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子計畫二：M4 無線網狀網路之存取控制協定及繞徑設計

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計畫參與人員： 鄭安凱、李奇育、張瑋倫、詹雯甄

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行政院國家科學委員會專題研究計畫期末報告

多天線多模多通道多速率無線網狀路—子計劃二：M^4無線網狀網路之存取控制協定及繞徑設計（1/3）

Media Access Control and Routing Protocols for M4 Mesh Networks

計畫編號：NSC　94－2219－E－009－005

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主持人：簡榮宏　　國立交通大學資訊工程系

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國立交通大學資訊工程系

中文摘要

無線網狀網路技術提升了無線網路存取範圍提升，而多重通道多重天線架構更可大幅增進網路效能。在本年度的計畫中，我們針對無線網路架構設計適用於多通道多天線環境的存取控制協定，使網路節點可於鏈結層中順利的傳遞訊息而不會造成過多的碰撞。為了進一步提升網路輸出，我們為通道使用方式進行最佳化的配置，設計可於效能和品質之間達到平衡的混合式通道指定演算法。此外我們以 VoIP 為應用，對 IEEE 802.11e 標準提出了跨層次的品質保證控制架構，使得 VoIP 有效地應用於此網路。

關鍵詞：無線網狀網路、存取控制，通道指定、服務品質保證，VoIP

Abstract

Wireless mesh network extends the wireless network's coverage. By equipping mesh nodes with multiple antennas, they can further utilize the benefits from multiple channels. In this project, we have investigated a media access control protocol for this multi-channel multi-antenna environment. It can avoid considerable collision at link layer. To further increasing the network throughput, we address the optimization issue and proposed approximating algorithms. Finally, the application of VoIP has been applied to this network. To achieve better performance, we designed a cross-layer control scheme for M^4 mesh networks with IEEE 802.11e interfaces.

Keywords: Wireless Mesh Network, Media Access Control, Channel Assignment, QoS, VoIP

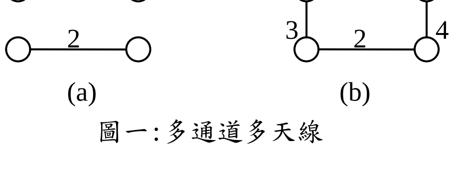
目錄

一、前言

近幾年，由於無線行動裝置的普及化以及對多媒體傳輸的需求，使得使無線寬頻技術的發展持續受到關注，而無線網狀網路(wireless mesh network)，則是預期將被整合進無線寬頻網路的重要技術之一。無線網狀網路是使用無線傳輸介面當網路主幹，因此不需要任何網路佈線等建設成本，有佈建快速及不受地形限制等優點，可將無線區域網路之涵蓋範圍快速的延伸至企業、校園以至於更大規模之無線都會網路(wireless metropolitan networks)。

雖然網路資源的存取變得便為便利，但在傳輸速率上仍然無法達到有線網路相同的水準，主要的原因就在於無線媒介的本質，無法將干擾隔絕於外，而正常的載波亦會影響鄰近節點的傳輸效能[1]，因此 IEEE802.11a/b/g 標準提供了多通道(multi-channels)的機制，使得鄰近的節點可透過非重疊的通道(non-overlapping)同時傳輸，避免彼此的干擾，進而提升整體網路的聚合頻寬(aggregated bandwidth)[2, 3]。這樣的技術，非常適用於節點間獨立運作的基礎架構型網路(infrastructure networks)[4]，但在隨意網路(ad hoc network)架構下，傳輸必需透過節點間相互的轉送(relay)方可達成，而相鄰的兩點亦需透過相同的通道方可收送，因此多重通道的使用必需在干擾程度(interference)與網路連接性(connectivity)上取得平衡[5]。而在無線網狀網路(wireless mesh network)架構中，網路更強調多點對多點(multipoint-to-multipoint) 同時傳輸的能力，因此每個節點都必需有能力協調來自四面八方的傳遞，而干擾模式與通道資源的利用也將變的更為複雜[6, 7]。

無線網路透過半雙工的天線於特定頻道上收送載波，而為了在減低干擾的同時，維持網路的連通性，節點可裝配數個天線(multi-antenna)[8, 9, 10]，如圖一(a)所示，上下兩對節點以不同的頻道傳輸彼此干擾，但卻斷為兩個子網路，若每個節點皆裝配兩個天線，圖一(b)，則可維持所有節點的暢通，而每個節點的 throughput 也可因此而倍增[5]。



圖一: 多通道多天線

此外在無線網路中，通道的位元錯誤率(bit error rate)會受到各種因素而變動，而且會隨著傳輸距離的增加而加劇[11, 12]，因此對鄰近來源端的節點，通常可以使用較高層級的變調機制(high level modulation scheme)來增加傳輸速率，反之則必需以較低層的變

調機制來處雜訊率(single noise rate)而降低速率[13]，以 802.11b為例，會因傳輸距離的不同有不同的傳輸速率，如 10m, 30m和 100m的距離內最高速率分別為 11Mbps, 5.5Mbps, 和 2 Mbps三種 [14]。除了 802.11b外，尚有a, g等標準，在不同的頻帶上以不同的速率、相容性、延遲、耗電量等提供多重的傳輸模式[15]，而目前的技術以可將多重模式以單一晶片和相同規格的天線，結合於無線設備上[16,17]。結合上述之特性、優點及可行性，本計畫以多天線、多通道、多速率、多模(multi-antenna multi-channel multi-rate multi-mode, M^4)作為無線網狀網路研究之基礎架構。

本研究為總計畫「多天線多模多通道多速率無線網狀路」之第二子計畫，共為期三年，主要探討M^4無線網狀網路之存取控制協定及繞徑設計。在第一年的進度中，我們針對無線網路架構設計適用於多通道多天線環境的存取控制協定(multi-channel multi-radio media access control)，使網路節點可於鏈結層(link layer)中順利的傳遞訊息而不會造成過多的碰撞(collision)。為了進一步提升網路輸出(throughput)，我們為通道使用方式進行最佳化的配置，設計可於效能和品質之間達到平衡的混合式通道指定演算法(hybrid channel assignment)。此外我們以VoIP為應用，對IEEE 802.11e標準提出了跨層次的品質保證控制架構(cross-layer QoS control scheme)，使得VoIP有效地應用於此網路。

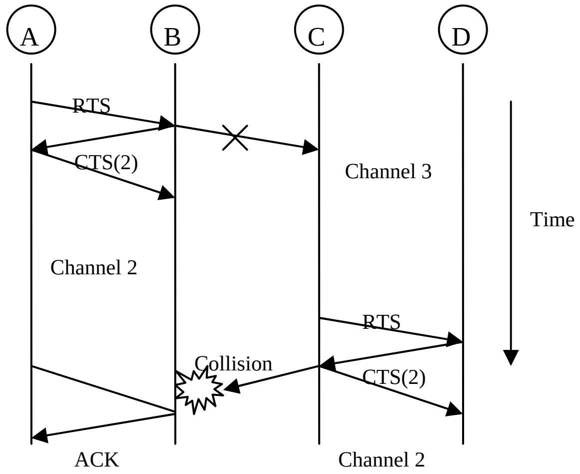
二、研究目的

為達到上述目標，本子計畫今年度已進行下列三項研究議題:

(1) 多通道多天線環境之存取控制協定(Multi-channel Multi-radio Media Access Control)

頻寬的共享需要載波偵測來避免碰撞，但無線網路的載波卻有距離上的限制，而造成隱藏節點的問題(hidden terminal problem)，因此 IEEE 802.11 MAC 協定利用分散式協調功能(distributed coordination function, DCF)中的 CTS/RTS 訊息交換，來達成虛擬載波偵測(virtual carrier sensing)的能力[18]。多通道的使用也是為了避免單一通道內的競爭，然而現有的分散式協調機制卻會在此環境下產生新的問題，稱作多重通道隱藏節點問題(multi-channel hidden terminal problem)[19]，以圖二來加以說明:主機 A 送出 RTS 要求傳資料送至主機 B，而主機 B 回覆 CTS 並指定使用通道 2 進行通訊，但此時主機 C 正利通道 3 與主機 D 通訊而無法偵測到主機 B 已使用了通道 2，因此當主機 C 重新要求與主機 D 傳送時，主機 D 就有可能指定 C 選擇通道 2 而產生碰撞。目前主要的解決方法可分為(1)跳頻模式 [20, 21]和(2)資料通道與控制通道分離[8, 9, 10, 22]以及(3)時間切割

[1]三種方式來達成，本計畫將綜合上述方法，發展新的改進方法，解決多重通道隱藏節點問題，新方法的主要考量則在於避免浪費寬使用率以及減少控制資料的延遲。



圖二: 多通道隱藏節點問題

(2) 通道指定之最佳化演算法 (Optimization on Channel Assignment)

在網線網狀網路中，當兩個節點要進行傳輸時，收送的兩端就必需處在一個共同的通道(common channel)才能行進溝通，而通道選擇的也會顯著的影響整體網路的效能，因此我們必需對共同通道的協調方式以及通道的指定方式進行設計，目前已有許多方法被提出，而這些方法大致可分為兩類 (a) 靜態通道指定(Static Channel Assignment): 在這種方式下，每個介面都會被指定一個永久使用的通道。由於通道是固定的，每個節點都能在網路初始時就知道鄰近通道指定的狀況，而不需在傳輸前耗費額外的封包來協調出共同的通道，然而固定的通道同時也意味著缺乏調整的彈性，每個介面只能對擁有相同通道的鄰近介面(neighboring interfaces)進行溝通，而當此通道壅塞時也無法切換至其他負載較輕的通導; (b) 動態通道指定(Dynamic Channel Assignment): 此方法與靜態正好相反，對通道對使用擁有最高的彈性，介面可依週遭或整體網路的使用狀況動態的調整所屬的通道，但相對的需要一個共同通道的協調機制(coordination mechanism)，當通道切換頻繁時，此機制無可避免將造成大量頻寬的浪費。

而我們探討的是一種混合通道指定(Hybrid Channel Assignment)方式，這種方式結合了靜態與動態的優點，不儘免除額外的協調機置，也同時擁有高度的調整彈性。我們的目標是對此通道指定方式設計最佳化以及近似的演算法，分析其理論上下界，並進一步以實測數據加以驗證。

(3) VoIP 於 IEEE 802.11e 之跨層次品質保證控制 (Cross-Layer Control Scheme for VoIP over 802.11e)

從 VoIP 技術發展以來，由於撥打 VoIP 電話的花費比一般電話節省許多，VoIP 漸漸成為一種殺手級的應用(Killer Application)，受到人們廣為使用。再隨著無線網路愈來愈普及，利用無線網路的可移動性，未來的人們不用固定坐在電腦桌前使用 VoIP 電話，而是可以像現在的行動電話這樣隨處講 VoIP 電話，我們稱之 Voice over Wireless LAN(VoWLAN)。

VoIP 這種即時性軟體對於 QoS 需求較嚴格，像是對於傳輸延遲(delay)、封包遺失率(packet loss rate)、延遲變動量(jitter)都有它的限制上限，例如傳輸延遲就不能超過 150 毫秒，超過的話通話品質差到無法容忍。然而無線網路有著頻寬低(lower bandwidth)、頻道變動 (channel variation) 大這 2 種限制。因為無線網路的頻寬低，如果同時網路上有大量的資料在傳輸，就會影響到 VoIP 的通話品質。而頻道變動更是使得 VoIP 處在不穩定的狀態下，也無法保證它的品質。因此要在無線網路上使用 VoIP 的服務，就必須比有線網路付出更大的努力克服這 2 項問題。我們的研究就是針對這二個限制，找出良好的解決方法以提升 VoIP 在無線區域網路中的通話品質。

三、文獻探討

下列分別對這靜態、動態和混合三種通道指定方法探討相關的文獻:

A. 靜態通道指定: 此方法是將某個頻道長時間分配給某個天線，為了可以同時使用多個頻道，每台主機必須配備和頻道數量一樣的天線，如此，將需要很高的硬體花費，這三篇文章[8][9][22]都是使用這個方法，不同的是如何決定在哪一個頻道做傳輸的方法，[8]以最近使用過的頻道為優先，若正在使用，則在閒置的頻道列中隨機選用一個，[9]則是選用最低功率的頻道，而且是由傳送者所決定，不同於[9]，[22]一樣選用最低功率的頻道，但是由接收者所決定。使用這個頻道給予的方法，雖然可以完全解決上面所述的兩個問題，但硬體的花費太高是主要的缺點。

B. 動態通道指定: 此方法是動態的分配某個頻道給某個天線，因此如何解決讓每個相鄰的主機可以互相找到對方是一個困難的問題，在這類方法中，都是使用時間同步的機制來解決這個問題，而[21][19][23][24]都是屬於這類方法，並且都是只使用一個天線，在[21]中，所有主機都有一個相同的頻道跳躍序列，而且必須隨著這個序列在同一時間作跳躍，當要傳輸封包之前，兩端必須實行接收端初始避免碰撞的交涉，然後留在目前的頻道做傳輸，其他的主機則隨著相同的序列繼續跳躍。[19]則是利用在 IEEE 802.11 節約能源模式中的 ATIM 窗戶，在此段時間裡傳輸控制封包，兩端交涉出要在那一個頻道上傳輸，在 ATIM 結束後，完成交涉的主機就切換到所決定的頻道做傳輸。[23]是將所

有的頻道分成兩類，一個是屬於控制頻道，其他則是屬於資料頻道，要到資料頻道做傳輸前，必須要在控制頻道先做交涉，當兩端選定頻道後，就切換到新的頻道，而為了避免隱藏終端問題，所以做傳輸之前必須等待一段時間，更新自己記載頻道上目前的狀況。在[24]中，每台主機各自擁有一個頻道跳躍序列，但在某個時間縫可以相遇的到，並且每台主機了解鄰近各主機的跳躍序列，所以也可以調整自己的跳躍序列與接收端重疊，而迅速地與接收端作傳輸。以上這些方法雖然只使用一個天線，但在多跳躍的網路環境裡，要維持時間同步是非常困難的。

C.混合的: 此方法兼具靜態和動態的給予頻道，所以屬於此方法的文獻都具備兩個以上的天線，這類文獻主要可分為兩種方法，第一種[10][25]是獨立出一個頻道作為控制頻道，然後將此頻道給予一個專為傳送控制封包的天線，另一個天線則是在其他頻道之間互相轉換，當控制天線在控制頻道上決定了傳輸的頻道，傳輸天線則會切換到這個頻道做傳輸。[26][27]則是給予每台主機一個固定的頻道，若其他鄰近的主機想要與其他主機做傳輸，則會切換到他的固定頻道，所以不會有會面的問題產生，而[26]改變傳送封包的功率，[27]在切換到一個新的頻道時等待一段時間，來解決多頻道隱藏終端的問題，這個方法可以增加頻道和天線的利用率，但這方法還是有些缺點，所以我們的方法是根據[27]的基礎架構，然後加以改進，由實驗得知我們網路的表現比他的方法好。

另一方面，IEEE 為了要提供支援服務品質(Quality of Service)的媒介存取控制協定(Media Access Control Protocol)，成立了工作小組 E(Work Group E)去修訂 IEEE 802.11 媒介存取控制協定以增加支援 QoS 功能，IEEE 稱這個新的協定為 IEEE 802.11e 媒介存取控制協定[28]。這個部份已在 2005 年 11 月完成，現在 802.11e 已是 IEEE 的標準了。

雖然藉由 802.11e 中的 EDCA 可有效給予不同種類的資料不同的優先權，然而簡單型的 802.11e 網路沒有使用機制對使用者做控管，會讓一些沒有申請 802.11e QoS 服務的使用者也使用到 802.11e 的服務(只要他有 802.11e 的無線網卡)，這對付費使用者是很不公平的。因此[29]這篇文章提出了一個基於策略(Policy-based)[33]的 802.11e 無線區域網路，在 802.11e 無線區域網路中加入一個伺服器—Wireless QoS Enhancer(WQE)。WQE 存放著使用者申請的服務等級契約(Service Level Agreement,SLA)，每當有一個新的資料流要通過無線基地台，無線基地台就會以一個客戶端(client)的角色向 WQE 詢問這個資料流是否有申請服務品質。如果有才會以 802.11e 的方法對這個資料流(flow)的封包設定存取類型(Acess Category)，否則就把它視為一般的資料流。但是這個架構有一個缺點，它的判斷是採取以資料流為單位的方式，這樣無線基地台要存的服務等級契約很多，對記憶體使用量限制高的無線基地台較不適合，所以我在後面會提出一個以行動工作站為

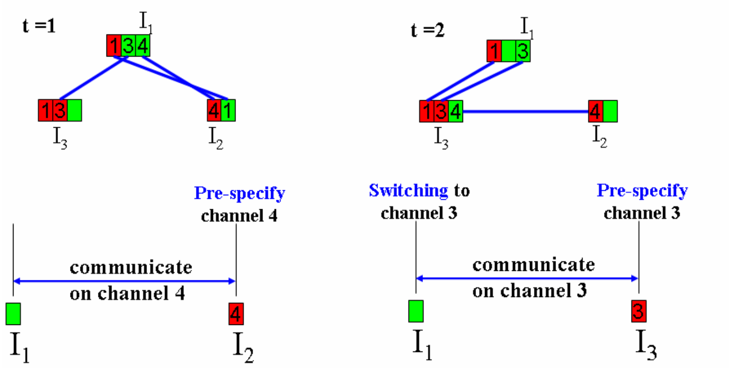
單位的方法，可以大幅減少無線基地台要存的使用者資料。

在參考論文[30]中提到 802.11e 無線區域網路為即時性(Real-time)應用程式(如 VoIP)提供了比 802.11 無線區域網路更好的 QoS 功能。但是當某個無線基地台同時服務很多運行高優先權應用程式(如 VoIP)的行動工作站時，便無法保證達到預期的服務品質效果[31]。特別是當其中某些行動工作站與無線基地台間的連結訊號弱或是受到干擾，就會對網路的效能及 VoIP 的通話品質產生不好的影響。這是因為當行動工作站與無線基地台之間的訊號變弱時，網卡會隨之自動選用資料傳輸速率較低的編碼方式來傳輸封包，希望藉此降低資料傳輸錯誤的機率。跟訊號強時所用高資料傳輸速率的編碼方式相比，傳輸同樣大小的封包，前者需要花費較久的時間，所以會耗費較多的網路資源。這對 VoIP 應用程式會有所影響，所以我們希望藉著加入可隨著網路狀況做調整的機制，來提升 VoIP 應用程式在 802.11e 無線網路中傳輸的通話品質。

在參考論文[32]中提到，使用可調變的(Adaptive) Codec – AMR，可改善 VoIP 的品質，所以我們可以利用跨階層調(Cross-layer)的概念設計一個可調變的 VoIP 應用程式(Adaptive VoIP Application)使它適用於環境多變的無線網路。

四、研究方法

在混合式通道指定法中，每個節點的介面可分為兩類，一類是靜態介面(Fixed Interface)，另一類是可轉換介面(Switchable interface)，當需要傳輸時，送端節點利用其可轉換介面進行傳送，而收端節點則利用其靜態介面進行接收，而所使用的通道則是靜態介可所決定。利用這個方法，節點並不需要在每一次的傳輸都先進行共同通道的協調，因為靜態介面所指定的通道是固定的，並不會雖時間而改變，因此只要在網路初始時廣播一次到其鄰近節點即可。由此可知此方法可免除大量的頻寬浪費，另一方面每個節點仍保有切換通道的彈性，只要利用其可轉換介面就可切換到所以鄰近靜態介面所指定的通道。如下例，紅色方塊為靜態介面，並固定於其所標示的通道上，而綠色為可轉換介面，在時間 t = 1 和 t = 2 時，所有的可轉介面都可直接切換到不同的通道，不需仍何協調機制。



圖三 :混合式通道指定法

底下(1)和(2)的部分，我們分別為此混合式通定法提出存取控制協定，以及介面和通道配置最化的演算法:

(1) 多通道多天線環境之存取控制協定(Multi-channel Multi-radio Media Access Control)

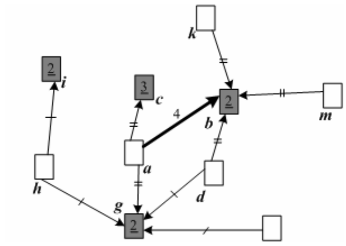
在[27]中，為了解決多頻道隱藏終端的問題，當可轉換介面切換通導後會原一段時間來更新 NAV。他們所利用的方法是等待一段最大封包的傳輸時間，然而此時間可能會過長而影響到整個網路的表現，所以我們利用碰撞的機率來計算出每台主機在每個頻道上需等待多少時間，而會產生碰撞主要是因為沒有接收到之前的控制封包，不知道鄰居的固定界面是否正在接收封包，所以我們以在頻道上鄰近的固定介面的個數，來計算出可能產生碰撞的機率，進而決定等待的時間。

因為每次轉換頻道時需要耗費時間，所以我們給予每個頻道設立一個行列，用行列來決定轉換頻道的時間，而不是利用每個封包來決定轉換頻道的時機，如果每次相鄰的封包都必須在不同的頻道上傳送，那傳送封包的資源花費是非常高的，且沒有效率，而在我們的協定上，利用轉換至新的頻道時停留一段時間，傳送屬於此頻道的行列裡的封包，來減少轉換頻道所帶來的資源花費。

(2) 通道指定之最佳化演算法 (Optimization on Channel Assignment)

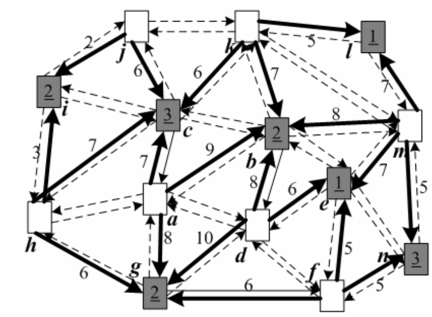
混合式通道指定包含兩大決策因素: A.介面種類選擇; B. 靜態介面通道指定。介面的種類選擇決定了網路的連通性(connectivity)，在實際的網路中每個介面可能有不同的傳輸半徑，因此若較大傳輸半徑的介面被選為靜態介面來進接收，則勢必會造成許多原本可以直接傳送的連通性，而一個節點的介面需有多少為靜態多少為可轉換，較多的可轉換介面可增加其傳送量，但相對的接收能力則會下降，反之亦然，因此因此介面種類的選擇同時需考量整體網路的拓撲和傳輸狀況。此外，介面種類的選擇也決定了主干擾(Primary Interference)的程度，由於一個介面同時只能進行一個收或送，當一對介面進行收送時，所以與其任一端直接連結的介面皆無法對其在進行收送，如圖四，當 a→b 傳

輸時，a→c、a→g、d→b、m→b、k→b 皆無法對 a 或 b 進行收送。另一方面，靜態介面通道的指定主要決定了次干擾(Second Interference)的程度，次干擾的產生是由於一塊區域面積內同時有多個傳輸使用相同的通道，這將造成彼此間的碰撞而降低網路的產出，如圖四，a→b、h→i、h→g、n→g、d→g 皆使用通道 2，因此同一時間內最只會有一對介面能成功傳輸。



圖四: 混合式通道指定之主干擾及次干擾

為了降低主干擾及次干擾所造成的影響而又能保有網路的連通性，我們必需對次兩決策因素進行最佳化的分析與設計。我們的目標是要使得每一個傳輸所可能造成平均的干擾最小化，如圖五，在此介面及通道配置下，平均每個傳輸會造成 6.36 個可能的干擾 (2+3+6+7+6+7+8+6+9+10+6+8 +7+5+5+6+8+5+7+7+5)/24。



圖五: 平均干擾

首先，我們利用整數線性規畫(Integer Linear Programming)將此問題公式化，使得此問題可被廣泛使用的簡化法(Simplex Method)所解決，如下線性函式:

$$\text{Maximize} \quad \sum_{(v,u)\in E} \sum_{i=1}^{K(v)} \sum_{j=1}^{K(u)} y_{i,j}(u,v)$$

Subject to

$$\sum_{(p,q)\in E_2(u,v)} \sum_{k=1}^{K(p)} \sum_{s=1}^{K(q)} H_{k,s,h}(p,q) y_{k,s}(p,q) \le 1,$$

$$\forall i = 1,2,\ldots,K(v),\ j = 1,2,\ldots,K(u), h = 1,2,\ldots,H, (v,u) \in E$$

$$\sum_{j=1}^{H} x_{i,j}(v) \le 1,\ \forall v \in V,\ \forall i = 1,2,\ldots,K(v)$$

$$\sum_{i=1}^{k_v} x_{i,j}(v) \le 1,\ \forall v \in V,\ \forall j = 1,2,\ldots,H$$

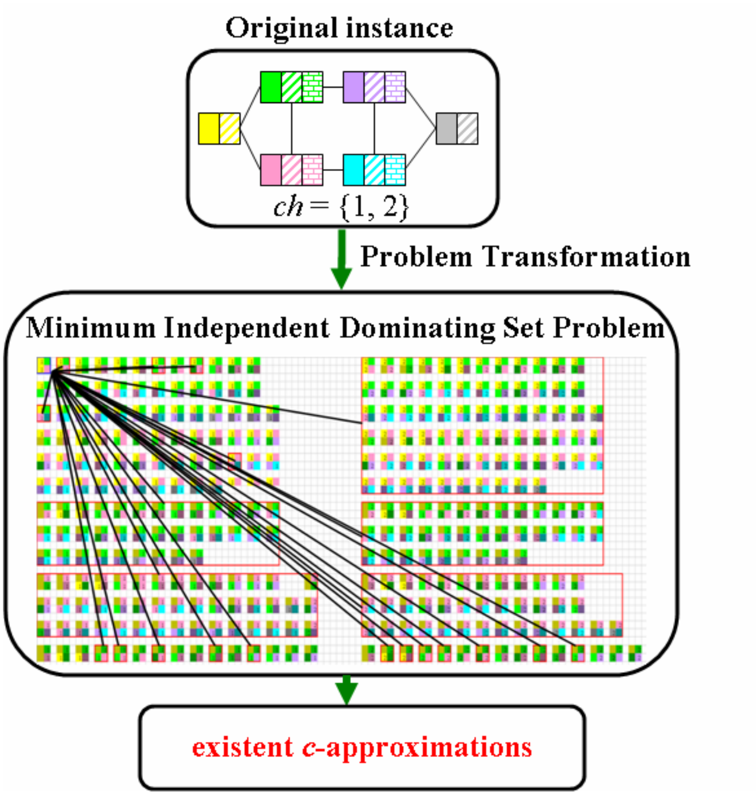
$$\sum_{i=1}^{K(v)}\sum_{j=1}^{H} x_{i,j}(v) > 0,\ \forall v \in V$$

$$\sum_{i=1}^{K(v)}\left(1-\sum_{j=1}^{H} x_{i,j}(v)\right) > 0,\ \forall, v \in V$$

$$x_{i,j}(v) = 0 \text{ or } 1,\ \forall i = 1,2,\ldots,K(v),\ j = 1,2,\ldots,H, v \in V$$

$$y_{i,j}(v,u) = 0 \text{ or } 1,\ \forall i = 1,2,\ldots,K(v),\ j = 1,2,\ldots,H, (v,u) \in E$$

為了更有效率地解決此問題，我們提出了具有常數上界的近似解演算法(c-Approximation)，主要的概念是用問題轉換的法式(Problem Transformation)，我們將此問題在多項式時間內轉換到最小獨立佔有集合(Minimum Independent Dominating Set Problem)問題上，使我們可以將現存的近似解演算法經過轉換後就可用來解決我們的問題，以目前最佳的解過，當網路拓撲的支度(degree)為B時，我們的問題可近似到B^2內。下圖為一示意圖，一個原本有五個節點的無線網狀網路可經由我們所提出的公式轉換成最小獨立佔有集合問題的輸入，因此我們可利用已存在的演算法經由反轉來近似原本問題的輸入。



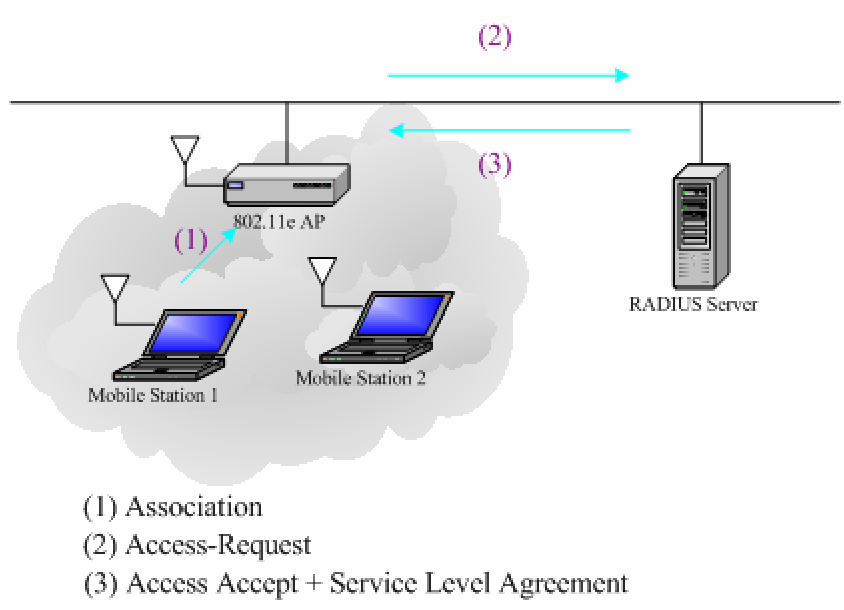
圖六: 問題轉換與近似解演算法

(3) VoIP 於 IEEE 802.11e 之跨層次品質保證控制 (Cross-Layer Control Scheme for VoIP over 802.11e)

此研究方法分為兩個部份，第一部份是改進 WQE[29]的方法，建構一個基於策略[33]的 802.11e 無線區域網路。第二部份提出一個跨階層調變 VoIP 應用程式(Cross-layer Adaptive VoIP application)來提升 VoIP 在 802.11e 無線區域網路運行的通話品質。

■ 第一部份 WQE[29]是以資料流為單位判斷是否符合策略規則(policy rule)，這種方法類似於 InterServ 的概念，它們的缺點就是需要儲存的策略資訊太多了，無法達到 scalable。所以我們設計了一個系統類似 DiffServ 的方法，以行動工作站的存取類別(Access Category)為單位，以減少無線基地台需要儲存的策略資訊數量。

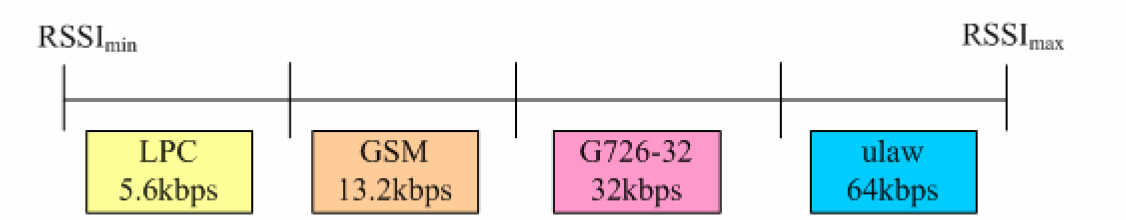
我們的方法如下圖所示:



圖七：主要架構

行動工作站連上無線基地台後，附在無線基地台上的 hostap daemon 會先幫工作站去和 RADIUS Server 做認證。認證完後，工作站申請的服務等級契約(Service Level Agreement,SLA)會附在 Access-Accept 中帶給無線基地台。我們是去修改 RADIUS Server，使之可以利用 RADIUS Message[34]中的 Vendor Specific Attribute 將服務等級契約傳給無線基地台，無線基地台收到服務等級契約後，每當有認證成功的行動工作站要傳送封包，無線基地台會去看 IP 封包的 DSCP 欄位的值判斷封包的類型，如果是 VoIP 的封包，無線基地台會再去服務等級契約中查看這個行動工作站是否有申請 AC_VO 的服務，如果有申請才將封包設成 AC_VO 的型態，使之擁有較高的傳送優先權。

■ 第二部份: 這個部份我們修改一個在 Window 平台上運行的開放原始碼 VoIP application—RAT(Robust Audio Tool)[9]。在 RAT 中新加入可根據行動工作站和無線基地台間訊號的強度值(RSSI(Receiver Signal Strength Indicator))來調整 VoIP codec 的方法。



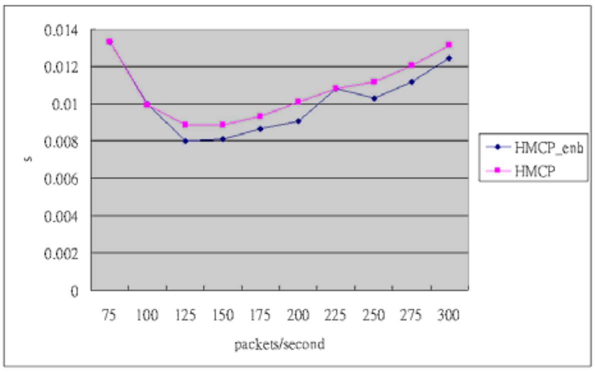
圖八： 選擇 codec 的方法

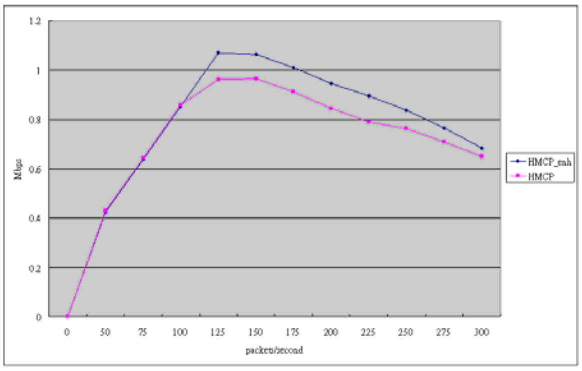
將$RSSI_{min}$~$RSSI_{max}$分為四個區段，每收集*n*個封包後就計算一次平均的RSSI值，若平均RSSI值落在某個區段A，就選擇用區段A對應到codec。

五、結果與討論(含結論與建議)

(1) 多通道多天線環境之存取控制協定(Multi-channel Multi-radio Media Access Control)

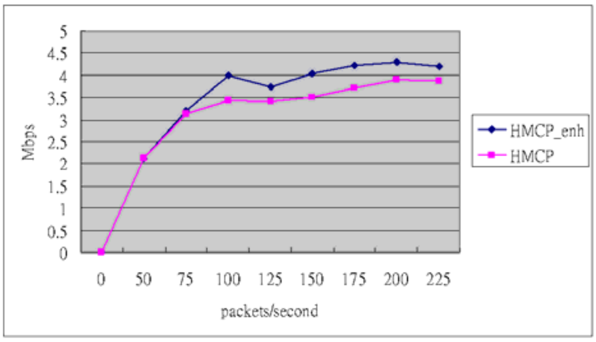
實驗平台式使用 ns2 來做模擬，圖 9 是在線型的拓撲上所做的實驗，總共有六個點，有兩條 UDP 流量，分別從兩端流向另一端，如此可測出隱藏終端的問題是否獲得紓解，圖 9 的橫軸為封包到達的速率，縱軸為網路的平均延遲，與[27]比較(HMCP)可發現我們的協定(HMCP_enb)在網路上的延遲是有改善的。圖 10 的橫軸是封包到達的速率，縱軸是網路的總輸出量，由圖可看出我們的總輸出量都比[27]改善許多。

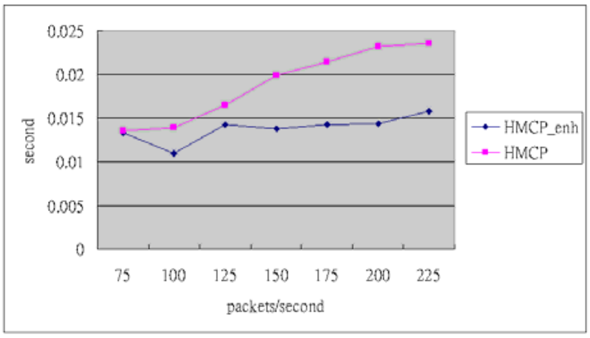




圖九：線型拓撲之平均延遲　　　　圖十：線型拓撲之總輸出量

圖 11 和圖 12 是在利用亂數所取得的拓撲上所做的實驗，總共有二十個點，有十條 UDP 流量，如此可測出我們的協定在實際狀況下是否真有改善，圖 11 的橫軸是封包到達的速率，縱軸是網路的平均延遲，與[27]做比較可發現我們在網路上的延遲是有改善的，且當封包到達的速率越快，所改善的幅度越大。





圖十一：隨機拓墣之平均延遲　　圖十二：隨機拓墣之總輸出量

(2) 通道指定之最佳化演算法 (Optimization on Channel Assignment)

此部分之成果主要以理論分析為主，當拓墣的支度最大為一常數B時，我們的演算法可近似到B^2 倍內，而當拓墣可被一等徑圓盤圖(unit disk graph)表示時，則可進似至 5 倍最佳解內。在未來的研究，我們將利用數據分析來驗驗理論的結果，並推導Relaxed ILP 線性規畫公式以作為數據分析時的理論下界。最後我們會把通道指定以及存取控制兩邊的結果加以結合，也就是以此通道指定演算法規畫網路初始時間的配置(包含介面種類決定以及靜態介面通道指定)，並以多通道多天線存取控制協定作為網路運作時的碰撞避免機制。

(3) VoIP 於 IEEE 802.11e 之跨層次品質保證控制 (Cross-Layer Control Scheme for VoIP over 802.11e)

下表中，我們分別測試 VoIP 應用程式在 802.11 與 802.11e 網路上運行的傳輸延遲，TCP 使用 Best Effort，而 VoIP 分別使用 Best Effort 和 Voice Priority，此外我們分別測試 5.3kbps 以及 64kbps 兩種 codec 傳輸的延遲:

表一: VoIP 於 802.11, 802.11e 以及不同 codec 大小之延遲

MAC	資料流優先權設定方式	64kbps	5.3kbps
802.11	VoIP：Best Effort TCP：Best Effort	140ms	104ms
802.11e	VoIP：Voice Priority TCP：Best Effort	11ms	9ms

(1)802.11e 與 802.11 效能比較

由橫的兩列來看 802.11e 與 802.11 效能比較，我們可以看到不管是使用資料傳輸率大的 codec 或是資料傳輸率小的 codec，VoIP 應用程式在 802.11e 網路上傳輸的延遲都會遠小於 802.11 網路，由此可以得知 802.11e 確實可以讓 VoIP 應用程式獲得較好的通

話品質。

(2)VoIP 應用程式 codec 效能比較

由第二列的資料來看，在 VoIP 應用程式在 802.11e 網路使用資料傳輸率小的 codec 可以減少傳輸延遲，這是因為使用資料傳輸率小的 codec 時，傳輸封包長度短，而且網路資料量也變少，所以傳輸延遲會變小。但是用資料傳輸率小的 codec 會使得聲音聽起來不是很清晰，所以我們不可以永遠都只用資料傳輸率小的 codec，我們的方法是將 802.11e 無線網路搭配跨階層調變的 VoIP 應用程式，讓 VoIP 應用程式隨著網路狀況調整 codec，以減小傳輸延遲，提升 VoIP 通話品質。這種方法在很多 VoIP 應用程式使用者位於訊號弱的狀況時,讓它們換用資料傳輸率小的 codec 可以減少對網路資源的浪費。

六、第二年工作計畫

無線網線透過繞徑機制將資料經由多重節點轉送至目地端，在實作中通道的指定會影響網路負載，繞徑的結果亦會改變個別虛擬連線的負載，因此指定的結果必需使多數的傳送競爭到足夠的頻寬，但頻寬的分佈又會進一步影響繞徑結果，而形成雙向相依的特性。另一方面，在對等式網路中，資料的繞徑經常透過尋找最少轉送點路徑來達成，以降低多重傳送所造成的延遲，可是在考量具有多重速率的環境時，雖然選取較遠的相鄰點的方法，可減少路徑轉送點的個數，但長距離的傳送意味著較低的傳輸速度，對持續性的傳輸未必能得到好處，因此繞徑的設計必需在延遲和速率之間取得平衡。而對於不同的模式，受到傳輸功率以及編碼模式的影響，距離與最大速度之遞減關係亦有不同的定義，此時若再使用不同的模式傳輸，將會再度提升繞徑設計之複雜度。因此，第二年我們將可先分別對多天線多通道無線網狀網路，以及多模多速率無線網狀網路，設計動態之繞徑演算法，並依此推導有用的性質、函式、和定理，作為M^4無線網狀網路動態繞徑研究的基礎。

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本年度本計畫衍生之論文如下：

附件一:A.-K. Jeng and R.-H. Jan, “Optimization on Hybrid Channel Assignment for Multi-channel Multi-radio Wireless Mesh Networks”, submit to *Globecom* 2006.

附件二:A.-K. Jeng and R.H Jan, The *r*-neighborhood Graph: A Novel Structure for Topology Control in Wireless Networks, accepted and to appear in *IEEE Transactions on Parallel and Distributed Systems*, 2005 (SCI).

附件三:Hsiao-Po Lin, Shih-Chang Huang and Rong-Hong Jan, "A Power-Saving Scheduling for Infrastructure-Mode 802.11 Wireless LANs," accepted and to appear in *Computer Communications*.

附件四:Yu-He Gau, Hung-Chi Chu, and Rong-Hong Jan, "A Weighted Multilateration Positioning Method for Wireless Sensor Networks" accepted and to appear in *International Journal of Pervasive Computing and Communications*, 2006.

Optimization on Hybrid Channel Assignment for Multi-channel Multi-radio Wireless Mesh Networks

Andy An-Kai Jeng and Rong-Hong Jan
Department of Computer Science
National Chiao Tung University
Hsinchu, 300, Taiwan, R.O.C

***Abstract*-The emergence of multi-channel multi-radio wireless mesh networks has given us many new opportunities and challenges. Particularly, the issue on how to appropriately assign channels to interfaces has gathered great importance in the recent publications. To efficiently utilize the channels diversity, the communicating channels should be negotiated between interfaces, which however would cause considerable overhead. To conquer this, a hybrid scheme was proposed to rule the way for assigning channels in literatures. In this paper, we formally model the hybrid scheme into an integer linear programming formulation. We provide the necessary as well as sufficient conditions of any feasible assignment which is not restricted to some specific goals. To for optimizing the linking-layer performance, we suggest a superior objective function which can minimize the potential interference of any transmission on average. We also show that a special case of the considered problem is APX-complete.**

I. INTRODUCTION

The wireless broadband technique has been continually grown in the last years. To extend the broadband access to the last mile, the deployment and research issues of the wireless mesh networks (WMNs) has attracted more and more attentions recently. In WMNs, a node can be either a mesh router or a mesh client [1]. By relaying packets thought multiple hops, the mesh clients, connecting to some mesh routers with AP functionality, are able to communicate to other nodes or access to the Internet.

Different from the traditional multi-hop networks, the WMNs put more stress on exploiting multiple non-overlapping channels to improve the network performance. By transmitting on different channels, the interferences among nodes can be greatly mitigated. However, for the time being, the available channels are still limited. In IEEE 802.11b/g (2.4GHz) and IEEE 802.11a (5GHz), at most 3 and 12 orthogonal channels are available only. Even in the further, wasting on the channels may not be a promising way. Therefore, the channel utilization should be properly spread among nodes to gain better spatial reusability. On the other hand, researches [1, 2, 5] have suggested that a mesh node should equip with more than one radio to increase its throughput in parallel. A node having multiple radios is able to transmit and receive packets simultaneously using non-overlapped channels.

In this paper, we consider the multi-channel multi-radio wireless mesh networks (MCMR-WMNs). Our goal is to assign channels on radios such that a link layer performance measurement can be optimized. A number of algorithms have been proposed to the channel assignment problem. According to Kyasanur and Vaidya [3], these algorithms can be categorized into threefold (scheme):

1) *Static Scheme*: In this scheme, the channels assigned on interfaces or links are fixed permanently or over a long period of time. The benefit of this scheme is that it requires no overhead for negotiating the communicating channel before transmission. All channels are decided in advance. However, as the traffic profile is changed, the bandwidth requirement on links may change. When this happens, the links or interfaces are not able to adequate to this change. Examples are [4, 5]. In [4], the channels assigned on the interfaces of each node are in the same pattern. *Das et al*. [5] provided a two optimization models for a link layer measurement. The channels are assigned on links rather than interfaces in [5].

2) *Dynamic Scheme*: This scheme allows the channels used by interfaces or links to be continually changed. However, to achieve this, communicating overhead for negotiating a common channel would be very heavy, if the traffic patterns are frequently changed.

3) *Hybrid Scheme*: This scheme is suggested by [3]. It combines the benefits of both static and dynamic schemes. In this scheme, an interface can be either fixed or switchable. For any fixed interface, a channel must be assigned on it in advance and can not be changed later. All transmission are initiated by switchable interfaces and received by fixed interfaces. When a switchable interface attempt to transmit to a target interface, it transmits using the channel assigned on the target interface. Because the channels of neighboring fixed interfaces of a switchable interface are always unchanged, the switchable interface can switch to the communicating channel immediately without any communicating overhead

In this paper, we propose an optimization model for the hybrid scheme. We formulate the channel assignment using hybrid scheme into an ILP formulation. A link layer performance measurement, named average potential interference per switchable link, will be given. The optimization models for the channel assignment problem with static and/or dynamic scheme for on MCMR-WMNs were discussed in [5, 6, 7, 8, 9]. Especially, [5] is most similar to our work. It concerns a link layer performance measure for the static scheme in the worst case. To the best of knowledge, there is no preview works concerning the optimization problem for the hybrid scheme.

II. NETWORK MODEL AND ASSUMPTIONS

In this section, we formally describe the networks under study. Some necessary assumptions and considerations are also explained. A given MCMR-WMN can be symbolically modeled as the following instance.

1) The mesh nodes correspond to a set V of n fixed nodes.

2) Each node v has $K(v)$ interfaces, where $K(v) \geq 2$. The set of interfaces equipped on v is represented by $I(v) = \{I_i(v) \mid i, =1, 2, \dots, K(v)\}$, where $I_i(v)$ is the i^{th} interface on v. Whether an interface is fixed or switchable is our decision variable either, rather than specifying in the instance. The rationale will be explained later.

3) Each interface $I_i(v)$ has its won transmission range, detonated by $R_i(v)$. The set of *neighboring nodes*, in the transmission range of $I_i(v)$ is denoted by $N_i(v)$. i.e. $N_i(v) = \{ u \mid u \in V, u$ is covered by $R_i(v)\}$. All neighboring nodes of v, reachable from some interface of v, is denoted by $N_i(v)$. i.e. $N(v) = \{ u \mid u \in N_i(v)$ for some $i\}$.(Note that $N(v)=\cup_{1\leq i\leq K(v)}N_i(v)$). In other words, by using different interfaces, the reachability of a node can be varied. This consideration is reasonable, since in the near further, multiple radios in different modes, like IEEE 802.11 a, b and g, can be equipped together on a single node to integrate heterogeneous networks. This will lead to unequal converge on each interface. Besides, even all interfaces are on the same mode, the topological control in MCMR architecture would result in diverse ranges on interfaces, which is beyond our scope.

4) Further, $NI_i(v)$ and $NI(v)$ are the sets of *neighboring interfaces* reachable from $I_i(v)$ and v, respectively. i.e $NI_i(v) = \{ I_j(u) \mid u \in N_i(v), j = 1, 2, \dots, K(u)\}$ and $NI(v) = \{ I_j(u) \mid I_j(u) \in NI_i(v)$, for some $i\}$.

5) The transmission range $R_i(v)$ of each interfaces $I_i(v)$ can be variant according the hardware feature or preliminary network planning. The neighboring interfaces $NI_i(v)$ of $I_i(v)$ is the set of interfaces which are within $R_i(v)$, not include $I_i(v)$ itself.

6) We assume for simplicity that all available channels are homogeneous to the entire deployment region. That is, H orthogonal channels are equally available to each node. To neglect trivial instance, we assume that $H \geq 2$. Actually, the assumption on the homogenous distribution of available channels can be released further. We will mention this in our discussion.

Then the linking status among nodes can be represented by a digraph $RG = (V, E)$, named *reachability graph* on nodes, where two nodes v and u has directed edge (v, u) in E if and only if u is reachable by some interface on v. i.e. $u \in N(v)$. Similarly, $RG_I = (V_I, E_I)$ is the reachability graph on interfaces, where $V_I = \{I_i(v) \mid v \in V, i = 1, 2, \dots, K(v)\}$, representing all interfaces in the networks, and an directed $(I_i(v), I_j(u))$ in E_I, if and only if $v = u$ or $I_j(u) \in NI_i(v)$. In this paper, we assume that any given instance of MCMR-WMN corresponding to RG is strongly connected. That is, for any two nodes v and u in V, there is some path in RG from v to u and vice versa[1]. Obviously, the connectivity of RG and RG_I are equivalent. In hybrid scheme, nodes should have at least on fixed interface for transmitting and at least one interface for receiving. Therefore, the assumption that $K(v) \geq 2, \forall v \in V$, is for the connectivity of instance either.

III. DESIGN PRINCIPLES

In the hybrid scheme, compared to the static scheme, the transmitting channels used by switchable interfaces can be dynamically chosen from those channels assigned on the neighboring interfaces which are fixed. So it is not necessary to determine a channel for any switchable interface before transmission. However, a new issue rises from this scheme: *whether an interface is fixed or switchable*? In the previous works [4], they generally assume that the roles of interfaces are given in advance. This assumption may however restrict the optimality within a specific consideration. We can treat the possible combinations of the transmitting channels used by all links in RG_I, including using no channel, as the whole solution space. Specifying the roles of interfaces in advance confines the solution space to a specific subgraph of RG_I, since those links, except between fixed and switchable, interfaces can be no longer in RG_I. Contrarily, unlike the dynamic scheme, the channels assigned on fixed interfaces are unchanged overtime, which means that for a fixed interface, it always uses the same channel to receive messages from different switchable interfaces. Therefore, we do not have to decide channels based on links. Instead, assigning a single channel on each fixed interface is representative enough. Accordingly, in this paper, the channel assignment problem on the hybrid scheme consists of the following two decisions

1) For each interface, decide whether it is fixed or switchable.

2) If an interface is fixed, assign a channel to it.

After an assignment is decided, a switchable interface can initiate an transmission with all neighboring interfaces which are fixed by switching to their channels. So, for any two interface $I_i(v)$ and $I_j(u)$, we say that a *switchable link* is from $I_i(v)$ to $I_j(u)$ if and only if the following three conditions are satisfied

1. $v \neq u$;
2. $I_j(u) \in N_i(v)$;
3. $I_i(v)$ is switchable and $I_i(v)$ is fixed.

By this, the switchability among interfaces of an assignment

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[1] The path from v to u and that from u to v are not necessarily the same.

can be represented by a digraph $SG_I = (V_I, E_{IS})$, named *switchability graph*, where E_{IS} corresponds to all switchable links between interfaces. Likewise, for an assignment, the corresponding connectivity of nodes, can be represented by a digraph $SG = (V, E_S)$, where a directed edge (u, v) is in E_S if and only if some switchable link is from u to v.

Given any instance, the feasibility of an assignment should be further constrained. In our work, any *feasible assignment* satisfies the following three considerations.

1) *About interfaces:* By definition, the channels assigned on fixed interfaces are always unchanged. Besides, the switchable interfaces are unnecessarily associated with any channel before transmission. So in a feasible assignment, any interface can be assigned at most one interface.

2) *About channels:* In MCMR architecture, multiple interfaces are available to exploit the usage of channels on separated spectrums simultaneously. Yet if two fixed interfaces are on the same channel as well as the same node, then there must be one of them useless, since they can never receive simultaneously at any time. In this case, for either of them, altering the channel or turning the role to switchable can never worsen.

Consequently, all channels assigned on the same node should be distinct. The constraint can be expressed as

3) *About connectivity:* Clearly, SG is always a subgraph of RG. Therefore, the assumption on the connectivity of RG is only a necessary condition for that of SG. We should further define the necessary as will sufficient condition for the connectivity (strongly connected) of any SG.

We say that a switchable link is *active* if it is used for transmission. For two reasons, a switchable link can not be active. First, the link is *blocked*, which means that the interfaces on either or both of the ends are occupied by other links for transmission. Obviously, any adjacent links can not be active simultaneously due to this reason. Second, the link is *interfered* by other transmissions. When a switchable link $(I_i(v), I_j(u))$ is active, the switchable links adjacent to the interfaces in $N_i(v)$ can not be active on the same channel of $I_j(u)$, since either transmitting or receiving of them can be interfered by $I_i(v)$'s transmitting (Note that some link could be blocked as well as interfered. In this case, we say the link is being blocked only, since no matter what channel is used by the link, it can not be active). See Figure 1 for example in which all interfaces transmit using channel 1: when the switchable link (a, b) is active , (a, c) and (g, b) are blocked, because their transmitting and receiving interfaces respectively are used by (a, b). On the other hand, (d, c), (d, f) and (e, c) can not be active either, since there are adjacent to c and/or d, which are within the transmission range of a.

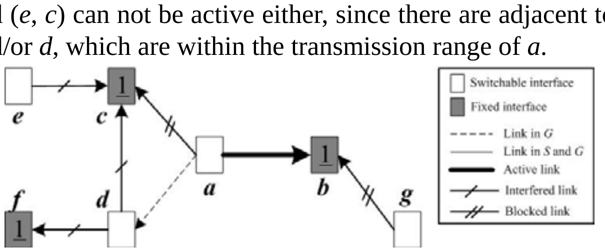


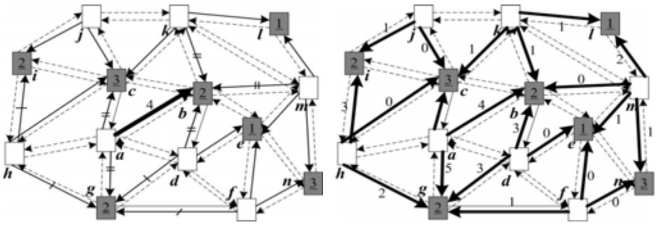
Figure 1: The links blocked or interfered by the transmission of (a, b)

In [5], they considered the static scheme. To optimize the link-layer throughput in the worst case traffic, in which all links contend for transmission at the same time, they set their goal to be maximizing the number of simultaneous transmissions between interfaces. Similarly, an institutive objective function for our consideration can be maximizing the number of switchable links which can be active simultaneously. However, this objective has many weaknesses. First, such goal would be more interesting for the hardest traffic profile, while ignore the most cases where only a pairs need transmit. Second, maximizing the simultaneous transmissions does not guarantee that any transmissions has the minimal interference to the other links A transmission causing high interference would lead to considerable degradation on throughput. Third, this objective does not take the channel diversity into consideration. In the hybrid scheme, a switchable interface can utilize varied channels to conquer dynamic traffic profiles by switching channels. Therefore, the higher switchability of interfaces could bring more channel diversity, where the switchability of an interface means the number of neighboring interfaces it can switch to. Third, this objective does not take the channel diversity into consideration. For any switchable interface, the *channel diversity* is the number of distinct channels it can switch to. An interface with higher channel diversity can be more adequate to dynamic traffic demands, since it has more choices to keep away from the intensively interfered channel.

From these viewpoints, in this paper, we aim to minimize an average-case linking-layer performance measure, named the *average potential interference per switchable link*, abbreviated as *APS*. Given an assignment, the *potential interference* of a switchable link $I_i(v)$ is the number of switchable links that are interfered by $I_i(v)$ on the same channel. Let us see Figure 2 (a): the potential interference of (a, b) is 4, including (h, i), (h, g), (d, g) and (f, g). Let TI be the total number of potential interferences and TS be the total number of switchable links. The *APS* is defined as

$$AIS = \frac{TI}{TS}. \tag{1}$$

This objective function provides an upper bound on the average number of interferences which would be caused by an arbitrary transmission. Figure 2 (b) illustrated a fully example, where totally has 21 switchable links and averagely no more than 1.19 transmissions (1.19 = 29 /21) could be interfered as a switchable interface is active.



(a) (b)

Figure 2: (a) the potential interference of a link (b) the potential interferences of all links.

IV. Problem formulation

The following, we are going to introduce the symbolic terms defining an assignment. Then, we discuss and formulize all essential constraints for characterizing a feasible assignment. Finally, an objective function dedicatedly tailored for the hybrid scheme will be presented.

A. *Decision variables*

First, we define the following binary variables to constitute an assignment based on interfaces:

$$x_{i,h}(v) = \begin{cases} 1 & \text{if channel } h \text{ is assigned on the } i^{th} \text{ interface of node } v \\ 0 & \text{Otherwise} \end{cases}$$

,where $v \in V$, $i = 1, 2, \dots, K(v)$ and $h = 1, 2, \dots, H$.

These variables are sufficient to model the two decisions. First, they dedicate the channels assign on interfaces. Secondary, an interface is switchable if there is no channel assigned on it, i.e i.e $\sum_{1\le h\le H} x_{i,h}(v) = 0$, otherwise, it must be fixed.

B. *Constraints*

The considerations of any feasible assignment have been verbally described previously. Now we formally define the sufficient and necessary conditions of the feasibility for any assignment.

Constraints on interfaces: For each interface, no matter being fixed or switchable, at most one channel can be assigned on it. Thus we have

$$\sum_{h=1}^{H} x_{i,h}(v) \le 1, \forall v \in V, \forall\, i = 1,2,\dots,K(v) \qquad (2)$$

Constraint on channels: To ensure that there is no channel assigned to any node for the second times, the constraint can be expressed as

$$\sum_{i=1}^{K(v)} x_{i,h}(v) \le 1, \forall v \in V,\ \forall h = 1,2,\dots,H \qquad (3)$$

Constraint on connectivity: In the hybrid scheme, a node can communicate with other nodes only if it has at least one fixed interface for receiving and has at least one switchable interface for transmitting. As a sequel, the obviously necessary conditions for connectivity can be given by

$$\sum_{i=1}^{K(v)} \sum_{h=1}^{H} x_{i,h}(v) > 0,\ \forall v \in V \qquad (4)$$

and

$$\sum_{i=1}^{K(v)} \left(1 - \sum_{h=1}^{H} x_{i,h}(v)\right) > 0,\ \forall v \in V \qquad (5)$$

Equations (3) and (4) indicate respectively that all nodes should have at least one fixed and switchable interfaces. However, they are not sufficient either. In *RG*, some neighboring node *u* of an interface $I_i(v)$ may be no longer reachable in an assignment if $I_i(v)$ is assigned to be fixed and no other interface of *v* can reach *u*.

Fortunately, when all transmission ranges on a node are equal (i.e. $\forall\ v \in V$, $R_i(v) = R_k(v)$, $\forall\ i, j \le K(v)$), equations (3) and (4) can be sufficient either: In the reachability graph *RG*, for any two nodes *v* and *u*, there is a path π from *v* to *u*. Considering any median node *w* and *w*'s next hop *t* on π, if the transmission ranges on *w* are all equal, *w* can reach *t* by any interface. Therefore, if some interface on *w* is switchable and some interface on *t* is fixed, *w* must be able to reach to *v*. Continuing the same argue on each *w*, the correctness follows.

C. *Objective function*

Now, we formulate our link-layer measure performance measurement *APS*, the *average potential interference per switchable link*. Let $r_{i,j}(u, v) = 1$ if $I_j(u) \in NI_i(u)$, 0, otherwise. A function $S_{i,j}(v,u)$ that indicates whether a switchable link is in *SG*, i.e. $S_{i,j}(v,u) = 1$ if $(I_i(v), I_j(u)) \in SG$, and $S_{i,j}(v,u) = 0$, otherwise, can be defined as follows.

$$S_{i,j}(v,u) = \left(1 - \sum_{h=1}^{H} x_{i,h}(v)\right)\left(\sum_{h=1}^{H} x_{i,h}(v) - \sum_{h=1}^{H} x_{j,h}(u)\right)^2 r_{i,j}(u,v) \qquad (6)$$

The first parenthesis returns 1 *i.f.f.* the two interfaces are not assigned the same role, the second returns 1 *i.f.f.* $I_i(v)$ is switchable, the last term indicates whether $I_j(u)$ can be reached by $I_i(v)$. So, the three conditions defining a switchable link is characterized by equation (6). By this, the total number of switchable links can be obtained by

$$TS = \sum_{v\in V} \sum_{i=1}^{K(v)} \sum_{u\in N(v)} \sum_{j=1}^{K(u)} S_{i,j}(v,u) \qquad (7)$$

Next, we consider the number of potential interference per link. Consider two pairs of interfaces $(I_i(v), I_j(u))$ and $(I_s(p), I_t(q))$. First, we define a function $F_{i,j,s,t}(u, v, p, q)$ to indicate whether they are switchable on the same channel. $F_{i,j,s,t}(u, v, p, q) = 1$, if it is, 0, otherwise. The function can be given as

$$F_{i,j,s,t}(v,u,p,q) = \sum_{h=1}^{H} \left(x_{j,h}(u) S_{i,j}(v,u)\right)\left(x_{t,h}(q) S_{s,t}(p,q)\right) \qquad (8)$$

Then, we need to identify whether $(I_s(p), I_t(q))$ is in the interference range of $(I_i(v), I_j(u))$. As shown by figures 1 and 2, $(I_s(p), I_t(q))$ can be interfered by $(I_i(v), I_j(u))$ only if either end of $(I_s(p), I_t(q))$ is reachable from $I_i(v)$ in *RG*. So, let $EI_i(v)$ denote the set of links in RG_I which are adjacent to some interface in $NI_i(v)$, all possible links which can be interfered by $(I_i(v), I_j(u))$ are in $EI_i(v)$. However, $(I_s(p), I_t(q))$ can be interfered, rather than be blocked, by $(I_i(v), I_j(u))$, only if either end is neither $I_i(v)$ nor $I_j(u)$. So, let $EI^-_i(v, u)$ be the subset of $EI_i(v)$ removing all links adjacent to neither $I_i(v)$ nor $I_j(u)$. The number of potential interference of $(I_i(v), I_j(u))$ can be defined as

$$TI_{i,}(u,v) = \sum_{(I_s(p), I_t(q)) \in EI_i^-(u,v)} F_{i,j,s,t}(v,u,p,q) \qquad (9)$$

, and the total number of potential interference links is

$$TI = \sum_{v\in V}\sum_{i=1}^{K(v)}\sum_{u\in N(v)}\sum_{j=1}^{K(u)} PI_{i,j}(v,u) \quad (10)$$

Finally, the average potential interference per switchable link can be obtained by

$$AIS = \frac{TI}{TS}$$

We name the channel assignment problem with the goal of minimizing *AIS* as *AISP*.

D. *The ILP formulation*

The final form of the ILP model is summarized below, which consists of totally $nH\sum_{v\in V}K(v)$ decision variables and $(H + 1)(n\sum_{v\in V}K(v) + 1) + 1$ constraints.

Minimize *AIS*

subject to

$$\sum_{h=1}^{H} x_{i,h}(v) \le 1, \forall v \in V, \forall\ i = 1,2,..., K(v)$$

$$\sum_{i=1}^{K(v)} x_{i,h}(v) \le 1, \forall v \in V,\ \forall h = 1,2,..., H$$

$$\sum_{i=1}^{K(v)}\sum_{h=1}^{H} x_{i,h}(v) > 0,\ \forall v \in V$$

$$\sum_{i=1}^{K(v)}\left(1 - \sum_{h=1}^{H} x_{i,h}(v)\right) > 0,\ \forall v \in V$$

$$x_{i,h}(v), \forall v \in V,\ \forall\ i = 1,2,..., K(v), \forall h = 1,2,..., H$$

Notice that in equation (1), the value of *TI* is determined by the two decisions, while the number of switchable links is solely determined by the roles of interfaces. Therefore, if the role of each interface is given in advance, we can concentrate on minimizing *TI*. We named this specialized problem as *TIP*.

IV. APPROXIMATION

In this section, we will show that in some reasonable consideration of network topology, *TIP* is APX-complete. In other words, in those cases, there is some polynomial-time algorithm such that the performance ratio of *TIP* can be bounded by a constant.

Given a graph *G*, a dominating set *D* of *G* is subgraph of vertices such that for each vertex *v* in *G*, *S* contains either *v* itself or some neighbor of *v* in *G*. An independent dominating set *T* of a graph is a dominating set such that no two vertices of *T* are connected an edge in *G*. The *minimum independent dominating set problem* (*MIDP*) is to minimize the vertices of *T* [13].

Let $P = (V, R, I, H)$ be an instance of *TIP*, where *R*, *I* and *H* are the sets of transmission ranges, interfaces and available channels, as defined in Section II, respectively. We show that there is a strict-reduction (*f*, *g*) from *TIP* to *MIDP*. That is, we shall prove the following

1) For every instance *P* in *TIP*, *f*(*P*) is an instance in *MIDP*.
2) For every feasible solution *T* to *f*(*P*), *g*(*T*) is a feasible solution to *P*.
3) *TI*(*g*(*T*)) – *TI*(*P*) = *MID*(*T*) – *MID*(*f*(*P*)), where *TI*(*g*(*T*)) (*MID*(*T*)) and *TI*(*P*) (*MID*(*f*(*P*))) are the result of *T* (*g*(*T*)) and optimal result of *P* (*f*(*P*)) respectively.

We now illustrate the transformation *f* from *P* to *f*(*P*)

(1) For any two interfaces $I_i(v)$ and $I_j(u)$ having a switchable link from $I_i(v)$ to $I_j(u)$ in *S*, we create *H* components $C_{i,j,h}(v, u)$, $h = 1, 2, \dots, H$.
(2) For any two components $C_{i,j,h}(v, u)$ and $C_{s,t,h}(p, q)$ having the same subscript *h*, if either $I_s(p)$ or $I_t(q)$ or both are within the $R_i(v)$, we have a node $N_{i,j,s,f,h}(v, u, p, q)$ in *f*(*P*).
(3) For any two nodes $N_{i_1,j_1,s_1,f_1,h_1}(v_1, u_1, p_1, q_1)$ and $N_{i_2,j_2,s_2,f_2,h_2}(v_2, u_2, p_2, q_2)$ in *f*(*P*), an edge is between then if and only if one of the following conditions is satisfied.
   i. $I_{j_1}(u_1) = I_{j_2}(u_2)$ and $h_1 \ne h_2$.
   ii. $I_{h_1}(q_1) = I_{h_2}(q_2)$ and $h_1 \ne h_2$.
   iii. $u_1 = u_2$, $j_1 \ne j_2$ and $h_1 = h_2$.
   iv. $q_1 = q_2$, $f_1 \ne f_2$ and $h_1 = h_2$.

Next, we show the function *g* which maps a independent dominating set *T* of *f*(*P*) to an assignment *g*(*T*) of *P*. An assignment is represented by binary matrix $y_{i,h}(v)$ such that $y_{i,h}(v) = 1$ means that the $I_i(v)$ is assigned channel *h*, and 0, otherwise. Then the function is defined as, for any $v \in V$, $i = 1,2,\dots, K(v)$ and $h = 1,2, \dots, H$, $y_{i,h}(v) = 1$ if there is some node $N_{i_k,j_k,s_k,f_k,h}(v_k, u_k, p_k, q_k)$ in *T* such that either $I_{j_k}(u_k) = I_i(v)$ or $I_{f_k}(q_k) = I_i(v)$, 0, otherwise.

Then, we can show that if *T* is feasible to *MIDP*, then g(*T*) is feasible to *TIP*. Compared to *AISP*, the conditions of defining the feasibility of *TIP* are equivalent, except for the connectivity, since this condition is determined solely be the roles of interfaces and these roles in *TIP* are given in advance. Besides, an additional condition for the feasibility is that all fixed interfaces are assigned at least one channel.

1) *Each fixed interface has at most one channel*: In *f*(*P*), if a node $N_{i_1,j_1,s_1,f_1,h_1}(v_1, u_1, p_1, q_1)$ is in *T*, which means that the two fixed interfaces $I_{j_k}(u_k)$ and $I_{f_k}(q_k)$ are assigned channel h_1 in *g*(*T*) , then any other node $N_{i_2,j_2,s_2,f_2,h_2}(v_2, u_2, p_2, q_2)$ such that $I_{j_1}(u_1) = I_{j_2}(u_2)$ and $h_1 \ne h_2$ or $I_{h_1}(q_1) = I_{h_2}(q_2)$ and $h_1 \ne h_2$ can not be in *T*, since an edge is between each of them an all nodes in *T* should be impendent.
2) *Each fixed interface has at least one channel*: It can see that the combinations of all fixed interfaces and channels are elaborated by nodes in *f*(*P*). Besides, by virtue of the dominating set, each node should be either in *T* or adjacent to some node *T*. Therefore, for each fixed $I_i(v)$, there must be some node in *T* with a subscript *h* consisting it. i.e. there must be some *h* such that $y_{i,h}(v) = 1$.
3) *All channels assigned on a node are distinct*: The conditions that $u_1 = u_2$, $j_1 \ne j_2$ and $h_1 = h_2$ ($q_1 = q_2$, $f_1 \ne f_2$ and $h_1 = h_2$) in above avoid that any two interfaces j_1 and j_2 on the same node u_1 (f_1 and f_2 on q_1) being assigned the same channel h_1.

Consequently, we can conclusion that any feasible solution of *f*(*P*) can be translated to a feasible assignment to *P* through a function g.

A node $N_{i,j,s,f,h}(v, u, p, q)$ in $f(P)$ corresponds to a interference that the activity on $(I_i(v), I_j(u))$ interferes the possible transmission on $(I_s(p), I_t(q))$. Thus $TI(g(T))$ is always equal to $MID(T)$. As we show above, for every feasible solution T to $f(P)$, $g(T)$ is a feasible solution to P. On the other hand, it can be easily evaluated that any feasible solution $g(T)$ to P can also be transformed to a feasible solution T to $f(P)$. Therefore, we have $g(T)$ is feasible if and only if T is feasible, which implies that that $TI(P)$ can be no worse than $MID(f(P))$ So, we can get that $TI(g(T)) – TI(P) = MID(T) – MID(f(P))$.

Halldorsson [10] show that for general n-vertex graphs, *MIDP* is not approximable within $n^{1-\varepsilon}$, for any $\varepsilon > 0$. Kann [11] show that *MIDP*, when restricted to graph of bounded degree, is APX-complete. Thus by the strict-reduction, if any instance P to *IPP* can be transformed to $f(P)$ having bounded degree, *TIP* is APX-complete. Moreover, we can observe that if all interfaces in P have bounded neighboring interfaces, the degrees in transformed instance re also bounded. So we can conclusion that *TIP* is APX-complete when considered bounded neighboring interfaces.

In wireless environment, the transmission power would exponential grown by distance. So, we usually hope to control the topology in advance so that each node covers a limited number neighboring nodes [12]. For this reason, the restricted consideration for P is quite reasonable.

Kann [11] also show that for d-regular graph, *MIDP* is approximable within $(d+1)/2$. As a sequel, *TIP* can also be approximable within $(d+1)/2$.

V. Conclusion

In this paper, we have studied the optimization model for the hybrid scheme of channel assignment problem on MCMR-WMNs. We have discussed and formulated the feasibility of an assignment into a set of necessary as well as sufficient constraints. To optimize the link layer performance, we have introduced an objective function *AIS*. A small *AIS* means that less interferences will be caused by an arbitrary transmission. We also show that *TIP*, a special case of *AISP*, is APX-complete for topology with bounded degree.

For the future research, it would be worth to investigate the optimization mode for the network layer performance measurement for the hybrid scheme. Besides, an efficient algorithm or distributed protocol for the hybrid channel assignment is attractive. In addition, whether the general problem *AISP* is also APX-complete or NPO-complete should be verified further.

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The *r* - Neighborhood Graph: An Adjustable Structure for Topology Control in Wireless Ad hoc Networks[#]

Andy An-Kai Jeng and Rong-Hong Jan[*]

Department of Computer and Information Science,

National Chiao Tung University,

1001 Ta Hsueh Road, Hsinchu, 300, Taiwan, R.O.C.

{andyjeng, rhjan}@cis.nctu.edu.tw

Abstract: In wireless ad hoc networks, constructing and maintaining a topology with lower nodes degrees is usually intended to mitigate excessive traffic load on wireless nodes. However, keeping lower nodes degrees often prevents nodes from choosing better routes that consume less energy. Therefore, a tradeoff is between the node degree and the energy efficiency. In this paper, an adjustable structure, named the *r*-neighborhood graph, is proposed to control the topology. This structure has the flexibility to be adjusted between the two objectives through a parameter r, $0 \le r \le 1$. More explicitly, for any set of n nodes, the maximum node degree and power stretch factor can be bounded from above by some decreasing and increasing functions of r, respectively. Specifically, the bounds can be constants in some ranges of r. Even more, the *r*-neighborhood graph is a general structure of both *RNG* and *GG*, two well-known structures in topology control. Compared with YG_k, another famous adjustable structure, our method can always results a connected planar with symmetric edges. To construct this structure, we investigate a localized algorithm, named *PLA*, consuming less transmitting power during construction and execute efficiently in $O(n\log n)$ time.

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[*] Corresponding author: Tel: +886-3-5712121 ext.31539.

Index Terms: Wireless ad hoc networks, topology control, energy-efficient, localized algorithm.

I. Introduction

Wireless ad hoc networks enhance the conventional deployment of communicating environments for many applications, such as conferences, hospitals, battlefields, search and rescue teams, etc. In these environments, the performance of network operations heavily depends upon the underlying topology [4]. For instance, the delivery rate would be significantly lower down as the underlying topology breaks. Therefore, appropriately controlling the topology is a crucial stage in communication. The topology control problem in wireless ad hoc networks has been widely studied in recent years [3, 15, 18, 19, 20, 23, 29, 32]. Generally speaking, the core of this problem is to determine set of wireless links such that the composed topology is able to achieve certain goals [23]. These goals would be variant depending upon the circumstances and could be either qualitative features or quantitative objectives. Since wireless nodes usually struggle with limited bandwidth and computation power, a genius way should be able to simultaneously achieve several goals. In this paper, we aim to control the topology with the following goals which are extremely desired in wireless environments.

1. *Symmetry*: The existence of asymmetric links may complicate many communication primitives. For instance, the MAC layer's ACK is hard to implement when some links are not bidirectional [21]. Besides, asymmetric links in topology would also cause inconsistent routing qualities at two ends.

2. *Connectivity*: Connectivity is unquestionably the most essential prerequisite in any communicable topology [23]. Two nodes *u* and *v* are strongly connected if there is a directed path from *u* to *v* and vice versa. A directed topology is strongly connected if all pairs of nodes are strongly connected. If the links are symmetric, we should aim at the *connectivity* of an undirected topology instead.

3. *Energy efficiency*: Energy is the most crucial resource in wireless nodes. Due to the severe path loss in radio carriers, transmitting with large ranges would exponentially run out of nodes' energy. Therefore, relaying messages through multiple hops with shorter ranges could usually consume less energy [24].How to choice the links between nodes for relaying is a critical point in this goal.

4. *Sparseness*: Numerous distributed and localized routing protocols are based on flooding [13]; however this may burden networks with unavoidable redundant messages. Thus, keeping a *sparse* topology, consisting of linear number of links [15], would be an ingenious way to shrink the expenditure from network operations.

5. *Maximum node degree*: For some nodes with overly-large degrees, the network flows will concentrate on them and rapidly draw out their energy. Besides, a larger node degree means tighter dependency among nodes, which is not expected when wireless nodes move frequently. Therefore, the maximum node degree over a topology should be bounded from above by some constant.

6. *Planarity*: A graph is *planar* if it has no crossed links inside. It is helpful for many geometric problems: The shortest path (least energy unicast route) can be quickly found in linear time when the underlying topology is planar [12]; Besides, in many position-based routing algorithms, the successful delivery can be guaranteed only if the underlying topology is a planar [2, 11].

Taking a further look, a topology having constant node degree must be sparse. So, we can concern the 5^{th} goal only. Unfortunately, keeping nodes with lower node degree would possibly sacrifice some potential links composing more energy efficient routes in topology. Therefore, empirically a tradeoff is between the node degree and energy efficiency [15]. For this reason, we aim to design an adjustable way so that the tradeoff can be adjusted flexibly.

In wireless ad hoc networks, due to the absence of a central arbitrator and the limited sensing range, centralized approaches [3, 30] are rarely attainable. Therefore, a variety of distributed approaches were proposed [17, 19, 29]. A distributed protocol passes messages hop-by-hop. This however may cause considerable overhead through the entire network. So, a localized approach is more preferred. According to Stojmenovic and Lin [27], a node using localized topology control method requires information within constant hop(s). However, in some localized approaches [15, 16, 18, 27], the operations should *recursively* depend upon the computed status or partial results from nearby nodes, which may hurt their practicability. Therefore, the following we define a new type of mythology for more practicability.

DEFINITION 1: An algorithm L is *purely localized* if it is localized and all operations depend upon only the information *inherent*[1] in nodes, available before any execution of L.

A purely localized topology control algorithm is more useful to large scale and high mobility environments, since the operation of a node is completely isolated from any execution of other nodes. Further, we say that a structure is *purely localizable* if we can construct it by a purely localized algorithm. Our goal is to investigate a purely localizable structure so that all desired goals listed above can achieve.

The rest of this paper is organized as follows. Section II specifies the network model and formally describes the problem under study. In Section III, we review and summarize the related works. The main structure, components, and their theoretical results are presented in Section IV. Some detailed derivations are explained in Appendix. In section V, we investigate an extended version of the main structure to comprehend our theoretical properties. In section VI, a purely localized algorithm is investigated to construct our structure. Finally, concluding remarks and some directions for the further research are given in the last section.

II. The Model and Problem

The wireless ad hoc network concerned in this paper consists of a set V of n wireless nodes distributed on two-dimension plane $\Re^2$. Each node is equipped with an omnidirectional antenna and can change its transmission range by adjusting the transmitting power at any level. The maximum transmission ranges are equal among all nodes. In other words, we can normalize the maximum transmission ranges of all nodes to be 1 for simplicity. In addition, each node u can obtain its position $P(u)$ through a lower-power GPS or some other ways [14], and an unique $id(u)$ is also available to each node u.

This network can be modeled as a *unit disk graph*, $UDG(V)$. In this graph, an edge uv exists if and only if the Euclidean distance between u and v, denoted as $||uv||$, is at most 1.

The least power required to transmit immediately between u and v is modeled as $||uv||^{\alpha}$, where α is typically taken on a value between 2 and 4, depending on the attenuation strength of the communication environment [5]. To measure the power efficiency of a topology, Li *et*

[1]The node's position and *id* are usually assumed to be inherited in nodes. See Section II for more explanation.

al. [15] defined a well-formed measure, named *power stretch factor.* We reintroduce it below. Let $\pi(u, v) = v_0v_1\ldots v_{h-1}v_h$ be a unicast path connecting nodes u and v, where $v_0 = u$ and $v_h = v$. The *total transmission power* consumed by path $\pi(u, v)$ is defined as

$$p(\pi(u,v)) = \sum_{i=1}^{h} \|v_{i-1}v_i\|^{\alpha} .$$

Let $\pi^*_{G(V)}(u,v)$ be the *least-energy path* connecting u and v in graph $G(V)$. Given a subgraph $G'(V)$ in $UDG(V)$, the power stretch factor of $G'(V)$ with respect to $UDG(V)$ is defined as

$$\rho(G(V)) = \max_{u,v\in V} \frac{p\left(\pi^*_{G'(V)}(u,v)\right)}{p\left(\pi^*_{UDG(V)}(u,v)\right)} .$$

On the other hand, the maximum node degree of graph $G(V)$ is defined as

$$d_{\max}(G(V)) = \max_{u\in V} d_{G(V)}(u) ,$$

where $d_{G(V)}(u)$ is the degree of node u in graph $G(V)$.

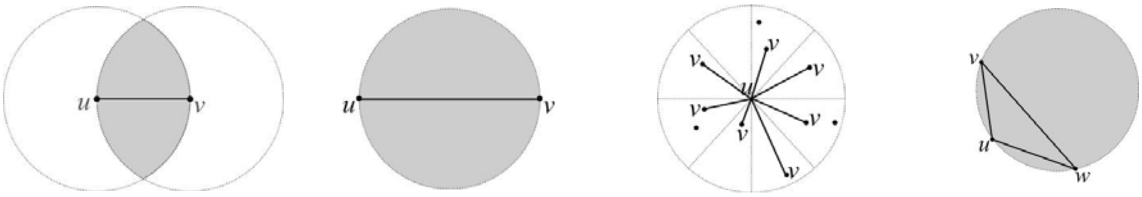
III. Related Work

Many localizable structures, used to control the network topology, have been proposed in literatures [15, 16, 18, 26], while only a few of them are purely localizable. The following we list four well-known structures. Most of them or their extensions are purely localizable:

- The *constrained Relative Neighborhood Graph* [28], denoted by $RNG(V)$, has an edge uv if and only if $\|uv\| \le 1$ and the intersection of two *open* disks[2] centered at u, v with radius $\|uv\|$ contains no node $w \in V$, see Fig. 1(a),
- The *constrained Gabriel Graph* [6], denoted by $GG(V)$, has an edge uv if and only if $\|uv\| \le 1$ and the open disk using $\|uv\|$ as diameter contains no node $w \in V$, see Fig. 1(b).
- The *constrained Yao Graph* [33] with a parameter $k \ge 6$, denoted by $\overrightarrow{YG_k}(V)$ is constructed as follows. For each node u, define k equal cones by k equal-separated rays originated at u. At each cone, a directed edge uv exists, if $\|uv\| \le 1$ and the cone contains no vertex $w \in V$ such that $\|uw\| < \|uv\|$. Ties are broken arbitrarily. $YG_k(V)$ is denoted as the underlying undirected graph of $\overrightarrow{YG_k}(V)$, see Fig. 1(c).
- A *Delaunay Triangulation*, denoted by $Del(V)$, is a triangulation of V in which the

[2] An open disk centered at point x with radius d is the collection of points with distance *less* than d from $P(x)$.

interior of the circumcircle of each Δuvw contains no node $w \in V$. The *unit Delaunay Triangulation*, denoted by $UDel(V)$, has all edges of $Del(V)$ except those longer than 1 [8, 18], see Fig. 1(d).



(a) (b) (c) (d)

Fig. 1: (a) $RNG(V)$ (b) $GG(V)$ (c) $YG_k(V)$, $k = 8$ (d) $UDel(V)$.

Let us discuss the properties of these structures and their extensions. We say a objective $f(.)$ of a structure $S(V)$ is *bounded* if there is a constant C such that $f(S(V)) \leq C$, for any set V of n nodes. Li *et al*. [15] showed that $d_{\max}(RNG(V))$ is unbounded if there is a node $u \in V$ having an unbounded number of neighbors adjacent to u at exactly the same distance in the underlying $UDG(V)$. To overcome this problem, Wattenhofer and Zollinger [32] proposed an algorithm to find a structure, denoted by $XTC(V)$. They showed that that $XTC(V)$ is a subgraph of $RNG(V)$ and the $d_{\max}(XTC(V))$ is at most 6. Especially, if there is no node having two or more neighbors at exactly the same distance in V, $XTC(V)$ is identical to $RNG(V)$ [24]. Their results infer the following theorem.

THEOREM 1: Given a set V of nodes on $\Re^2$, if there is no node having two or more neighbors at exactly the same distance, then $d_{\max}(RNG(V)) \leq 6$. □

We denote the condition in Theorem 1 as assumption *AS*. That is,

AS : There is no node in V having two or more neighbors at exactly the same distance.

This theorem reveals that even $RNG(V)$ has no constant bound on its node degree, it is still useful since the distances of nodes in real world are rarely exactly the same. The constrained Gabriel Graph $GG(V)$ has the least power stretch factor 1, in comparison with the unbounded power stretch factor $n - 1$ of $RNG(V)$ [15]. However, $d_{\max}(GG(V))$ could be as large as $n - 1$. An extended structure, *Enclosure graph* [16, 14, 24], denoted by $EG(V)$ is generalized from $GG(V)$. It can always result a subgraph of $GG(V)$ [16]. Even so, its maximum node degree is still unbounded [20, 24].

To overcome the tradeoff between the maximum node degree and the power stretch factor, an adjustable structure, having the flexibility to be adjusted between the two objectives, becomes more attractive. $\overrightarrow{YG}_k(V)$ is an adjustable structure. It can be adjusted through a parameter k such that for any given k, the maximum *out*-degree is at most k, and the power stretch factor is at most $1/\left(1-(2\sin\pi/k)^{\alpha}\right)$ [15]. We say an objective $f(.)$ of an adjustable structure $S_k(V)$ with parameter k is *partially bounded* if there is *at least one* k_0 such that $f(S_{k_0}(V))$ is *bounded*. According this definition, the maximum *out*-degree and power stretch factor of $\overrightarrow{YG}_k(V)$ are partially bounded since for some ranges of k, k and $1/\left(1-(2\sin\pi/k)^{\alpha}\right)$ are constants. However, the asymmetric edges of $\overrightarrow{YG}_k(V)$ may lead to large *in*-degrees even when k is very small [15]. So, $d_{\max}(YG_k(V))$ can be neither bounded nor partially bounded. To improve this, an extension of $\overrightarrow{YG}_k(V)$, named *Yao and Sink*, was proposed [15, 17, 29]. It can limit the maximum node degree in $(k+1)^2-1$ and result symmetric edges. Unfortunately, in this structure the neighbors of some node should be recursively determined by one another so that it can not be purely localizable. The unit Delaunay triangulation $UDel(V)$ has bounded power stretch factor. However, neither $Del(V)$ nor $UDel(V)$ can be computed locally. So, Li *et al*. [18] suggested a localized version of the Delaunay graph, denoted by $LDel^{(h)}(V)$, where h means that each node uses at most k-hop information. The power stretch factor of $LDel^{(k)}(V)$ is bounded for all $k \geq 1$. Even so, its maximum node degree is not bounded for any h.

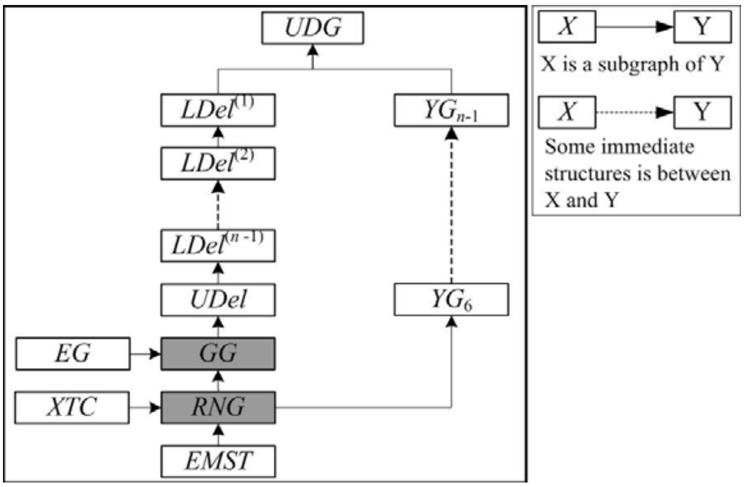


Fig. 2: The relations of the pure localizable structures and their extensions.

The relations among these structures were studied in several papers [7, 10, 16, 22, 24, 33]. We summarize them on Fig. 2, where $EMST(V)$ is the Euclidean minimum spanning tree of $UDG(V)$. With these relations, their connectivity can planarity can be easily inferred.

Regarding the connectivity: we know that $EMST(V)$ is connected if $UDG(V)$ is itself a connected component of V. Therefore, when $UDG(V)$ is connected, all graph containing $EMST(V)$ are connected. That is, $RNG(V)$, $GG(V)$, $EG(V)$, $UDel(V)$, $LDe^{(k)}l(V)$, $YG_k(V)$ are all connected. The connectivity of $XTC(X)$ was proven by different way [24].

Regarding the planarity: $LDel^{(k)}(V)$ is planar for any $k \geq 2$ [18]. Therefore, all subgraphs of $LDel^{(2)}(V)$ are planar. That is, $UDel(V)$, $GG(V)$, $EG(V)$, $RNG(V)$, $XTC(V)$, $EMST(V)$ are all planar. On the contrary, $\overrightarrow{YG}_k(V)$, and $LDe^{(1)}l(V)$ can not avoid producing the crossed link, so they are not planar [15, 18]. Table 1 summarizes above discussion.

Table 1: The properties of the four main purely localizable structures.

	Power stretch factor	Maximum node degree	Planar	Symmetric	Connected
$RNG(V)$	Unbounded	Bounded (with *AS*) Unbounded (without *AS*)	Yes	Yes	Yes
$GG(V)$	Bounded	Unbounded	Yes	Yes	Yes
$YG_k(V)$	Partially bounded	Unbounded	No	No	Yes
$LDel^{(k)}(V)$	Partially bounded	Unbounded	No ($k=1$) Yes ($k \geq 2$)	Yes	Yes

From above table, we can see that no presented structure can bound or even partially bound the two objectives. Besides to the best of our knowledge, no other structure can be purely localizable and achieve this goal. Therefore, we will propose the first purely localizable structure, named *r-Neighborhood Graph*, to fill this gap. This structure is adjustable and can always result connected planar with symmetric edges. In addition, we can show that our structure is a generation of both $GG(V)$ and $RNG(V)$.

Apart from the purely localizable structures, several composite methods, based on combining two or more existent structures, were investigated in the last few years [17, 19, 25, 31]. Conceptually, the main idea is to use the virtue of one structure to patch up the fault in the other structures. For examples, the *ordered Yao structure*, denoted as $OrdYao(V)$ [1], is a variation of $YG_k^*(V)$. It has the partially bounded maximum node degree and length stretch factor. However, the planarity can not be guaranteed. Therefore, Wang and Li [19, 31] applied $OrdYao(V)$ onto $LDel^{(2)}(V)$ to avoid the crossed edges produced by $OrdYao(V)$; Song *et al*. [25]

improves it by applying the *OrdYao*(V) on *GG*(V), using only one-hop information. However, the construction of *OrdYao*(Y) requires exchanging the computed status as well as partial results between nodes. Consequently, none of them is purely localized or purely localizable.

IV. The *r* –Neighborhood Graph

In this section, we introduce a new adjustable structure. First, we define a region on $\Re^2$. It will be used to compose our structure. Let x be any point on $\Re^2$, the *open disk* and *circle* centered at $P(x)$ with radius d are denoted as $D(x,d)$ and $C(x, d)$, respectively. The region is defined as follows.

DEFINITION 2: Given a nodes pair (u, v) on $\Re^2$, the *r-neighborhood region* of (u, v), denoted as $NR_r(u, v)$, is defined as:

$$NR_r(u,v) = D(u,\|uv\|) \cap D(v,\|uv\|) \cap D(m_{uv}, l_{uv}),$$

where m_{uv} is the middle point on uv, $l_{uv} = (\|uv\|/2)(1 + 2r^2)^{1/2}$, and $0 \le r \le 1$.

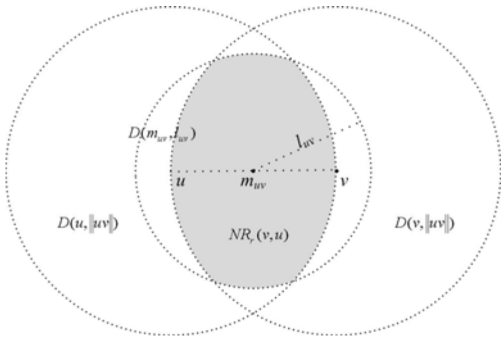


Fig. 3: The r-neighborhood region of nodes u and v.

When no confused, we use m and l instead of m_{uv} and l_{uv} respectively. In Fig. 3, the shaded region intersected by the three open disks sketches an example of the r-neighborhood region. This region is obviously equivalent to the following point set:

$$NR_r(u,v) = \{P(x) \in \Re^2 \mid \|ux\| < \|uv\|, \|vx\| < \|uv\|, \|mx\| < l\} \qquad \text{Eq. 1}$$

For any node w located on $NR_r(u, v)$, this region limits the power consumed by path uwv. This property is shown in Lemma 2 and derived in Appendix.

LEMMA 2: Given two nodes u and v on $\Re^2$, for any node w such that $P(w) \in NR_r(u, v)$, $p(uwv) < \|uv\|^{\alpha}(2 + r^{\alpha})$, for all $\alpha \ge 2$.

This lemma explains why we call such plane a *neighborhood region*: For any node w located in the region $NR_r(u, v)$, it should be an alternative neighbor for u with respect to v, in the sense that the power required for relaying from u to v through w is no greater than $1 + r^{\alpha}$ times of the immediate transmission. Based on this region, we structure is defined below.

DEFINITION 3: Given a set V of nodes on $\Re^2$, the *r-neighborhood graph* of V, denoted as $NG_r(V)$, has of an edge uv if and only if $\|uv\| \le 1$ and $NR_r(u, v)$ contains no node $w \in V$, where $0 \le r \le 1$.

By Definition 3, if edge uv is not in $UDG(V)$ or a node w is inside $NR_r(u, v)$, there is no direct link connecting u and v in $NG_r(V)$, which mean that all transmissions between u and v should be relied through some other node(s) in $NG_r(V)$. Now, we explore the desired properties in our structure. Before this, we shall discussion the following relations.

LEMMA 3: For any set V of nodes on $\Re^2$, $RNG(V) \subseteq NG_r(V) \subseteq GG(V)$, for all $0 \le r \le 1$.

Proof. Consider the open disk $D(m, \|uv\|/2)$, defining $GG(V)$. Suppose $uv \in NG_r(V)$, the region $NR_r(u, v)$ has no node inside. Since $D(m, \|uv\|/2)$ is obviously a subregion of $NR_r(u, v)$, for any $0 \le r \le 1$, there is also no node in $D(m, \|uv\|/2)$. Therefore, according to the definition of $GG(V)$, we get $uv \in GG(V)$. On the other hand, consider the two open disks $D(u, \|uv\|)$ and $D(v, \|uv\|)$, defining $RNG(V)$. Suppose $uv \in RNG(V)$, no node is inside the intersection of $D(u, \|uv\|)$ and $D(v, \|uv\|)$, which obviously covers the region $NR_r(u, v)$, for any $0 \le r \le 1$. Therefore, no node can be inside $NR_r(u, v)$ and we get $uv \in NG_r(V)$. □

Specifically, as $r = 0$, $NR_0(u, v) \equiv D(m, \|uv\|/2)$, which is the disk defining $GG(V)$. On the contrary, as $r = 1$, $NR_1(u, v) \equiv D(m, \|uv\|)$, which is the disk defining $RNG(V)$. Therefore, $GG(V) \equiv NG_0(V)$ and $RNG(V) \equiv NG_1(V)$. So, we can conclude the following theorem.

THEOREM 2: The r-neighborhood graph is a generalized structure of both the restricted Gabriel graph and the restricted relative neighborhood graph.

Since a subgraph of a planar graph is always planar, and a supergraph of a connected graph is always connected, with the planarity of $GG(V)$ and connectivity of $RNG(V)$, we can infer the following two theorems.

THEOREM 3: For any set V of nodes on $\Re^2$, $NG_r(V)$ is planar, for all $0 \le r \le 1$.

THEOREM 4: For any set V of nodes on $\Re^2$, if the underlying $UDG(V)$ is connected, $NG_r(V)$ is *connected*, for all $0 \le r \le 1$.

Now we consider the energy efficiency and node degree of $NG_r(V)$. We will show that the upper bound of $\rho(NG_r(V))$ is increased by r and contrarily the upper bound of $d_{\max}(NG_r(\mathrm{V}))$ is decreased by r. In other words, the r-neighborhood graph is adjustable to the two objectives through the parameter r. With these results, we can further show that the power stretch factor and maximum node degree are partially bounded in our structure. Before these, a property proposed by Li *et al.*[15] shall be mentioned first. It can be used to simplify our proof.

LEMMA 4 [15]: Given a subgraph $G'(V) \subseteq UDG(V)$ and a constant C, $\rho(G'(V)) \leq C$ if and only if for any edge uv in $G(V)$, there is a path $\pi(u, v)$ in $G'(V)$ such that $p_{G'(V)}(u,v) \leq C\|uv\|^{\alpha}$. This lemma indicates that to derive an upper bound for $\rho(NG_r(V))$, it is sufficient to the consider only those nodes pairs having direct links in $UDG(V)$. So, we aim to derive a strictly decreasing function $F(r)$, such that for any uv in $UDG(V)$, a path $\pi(u,v)$ is in $NG_r(V)$ such that $p(\pi(u,v)) \leq F(r)\|uv\|^{\alpha}$. To achieve this, we investigate an algorithm **EXPANSION** with an input of any two nodes (u, v) and outputs subgraph S of $NG_r(V)$ related to (u, v). Let $P(S)$ be the total transmission power of edges in S. i.e. $P(S) = \sum_{st \in S} p(s, t)$. We can show that there is some path in S connecting (u, v) and $P(S) \leq F(r)\|uv\|^{\alpha}$.

ALGORITHM EXPANSION

Input: A nodes pair (u, v) in V

Output: A subgraph S and a positive value P.

Step 1: $S = \{\}$, $S' = \{(u, v)\}$, $Q = \{u, v\}$, $P = \|uv\|^{\alpha}$;

Step 2: When some node pair (s, t) is in S such that a node $w \in NR_r(s, t)$

    $S' = S' - (s, t)$;

    If $w \notin Q$ then

        $S' = S' \cup (s, w) \cup (w, t)$;

        $Q = Q \cup \{w\}$;

        $P = P + (\|st\|r)^{\alpha}$;

    Otherwise,

        $S' = S' \cup (s, w)$;

Step 3: $S = \{xy \in NG_r(V) \mid (x, y) \in S'\}$;

Step 4: Stop and output E and P.

In this algorithm, S' is a set of nodes pairs, in which an edge st in $NG_r(V)$ can be a part of S only if its two ends (s, t) are in S' as described at step 3. So, to determine S, we have discuss the S' first. Initially, S' contains only (u, v). Then, it will be recursively expanded as follows: for each (s, t) in S', if a node w is in $NR_r(s, t)$ and not considered before, replace (s, t) with $(s,$

w) and (w, t); if a node w is in $NR_r(s, t)$ but considered before, replace (s, t) with (s, w); Otherwise, keep (s, t) unchanged. We use the set Q to record the considered nodes.

When some (s, t) is in S' such that a node $w \in NR_r(s, t)$, no matter w is considered or not, by Eq. 1, the replaced nodes pair(s) must be shorter than ||st||. i.e. $\|sw\| < \|st\|$ and $\|wt\| < \|st\|$. Thus after finite iterations, each node pair in S' can be replaced by another node pair with shortest distance. So, the algorithm is terminable. Now we show that (u, v) is connected by some path in the subgraph S when termination.

LEMMA 5: Given any set V of node on $\Re^2$, for any two nodes u and v in V, if edge uv is in $UDG(V)$ and $UDG(V)$ is connected, there is some path in S connecting (u, v).

Proof: Since Q includes u and v, we can prove this lemma by showing that all nodes in the Q are connected in S. For each expansion of S', we define a dummy graph S" in which an edge st exists if and only if (s, t) is in S' (Note that any edge in S" is not necessarily in either $UDG(V)$ or $NG_r(V)$). First, we show that at any iteration, all considered nodes in Q are connected by S". Initially, Q is connected by S", since $S' = \{(u, v)\}$ and $Q = \{u, v\}$. We assume for induction that all nodes in Q are connected by S" at k-th iteration. Then, we show that it is true for the next iteration. At k+1-th iteration, if there is no pair in S' satisfies the entrance condition of step 2, the claim is correct, since Q and S" are unchanged; Otherwise, a node pair $(s, t) \in S'$ is expended. In this case, if the chosen $w \notin Q$, w is connected with all nodes in Q via dummy edges sw and wt; otherwise, $w \in Q$, which implies all nodes in Q are still connected by S" as the previous iteration. As described above, the distance of any expended nodes pair is no longer than the previous one. So, if uv is in $UDG(V)$, all edges in S" are also in $UDG(V)$. Then, as the algorithm processes to step 3, no nodes can be in the r-neighborhood region of any nodes pair in S'. With these two facts, all dummy edges in S" are also in $NG_r(V)$ when termination. So S is equivalent to the last S". Consequently, if $UDG(V)$ is connected, by Theorem 4, all nodes in the last Q are connected S. □

Then we derive a strictly decreasing function $F(r)$ using the value P in this algorithm.

LEMMA 6: Given any set V of n nodes on $\Re^2$, for any two nodes u and v in V,

$$P(S) \leq F(r)\|uv\|^{\alpha} \quad \text{and} \quad F(r) = 1 + (n-2)r^{\alpha}$$

for all $0 \leq r \leq 1$ and $\alpha \geq 2$.

Proof: Let $P(S') = \sum_{(s,t)\in S'} p(s, t)$. We show that $P(S') \le P$ at each iteration of step 2. Initially, $S' = \{(u,v)\}$. We can get $P(S') = \|uv\|^{\alpha} = P$. Then at the first iteration, if no node w is in $NR_r(u, v)$, the claim remains true since neither P nor S is changed; Otherwise, a node w is in $NR_r(u, v)$. Besides, any chosen w can not be in Q, since no nodes except u and v are in Q so far. So, uv is replaced by vw and wv. By Lemma 2, $P(vw)+P(wu) \le P(uv)(1+r^{\alpha}) = P + (\|uv\|r)^{\alpha}$. Consequently the new P remains a upper bound of $P(S')$. We assume for induction that $P(S') \le P$ at k-th iteration. Then we prove the claim is true at the next iteration. If the entrance condition of step 2 is not satisfied or the chosen $w \notin Q$, it can be proved by the same reasons as in the first iteration. Otherwise, assume (s, t) is taken, st is replaced by only sw. By Eq. 1, $P(vw) \le P(uv)$, which implies that the unchanged P is still an upper bound of $P(S')$. Besides, Eq. 1 further implies that all distance of two nodes in E are no greater than $\|uv\|$. So, another upper bound P' can be get by replacing $P = P + (\|st\|r)^{\alpha}$ by $P' = P' + (\|uv\|r)^{\alpha}$. Moreover, we can observe that the situation that as a w is chosen from some $NG_r(s, t)$ is not in Q never happens over $n - 2$ times, since in this case the size of Q must be increased 1. Consequently, $P(S') \le P \le P' \le P(uv) + P(uv)r^{\alpha}(n-2)$. Finally, we get $F(r) = (1+r^{\alpha})(n-2)$. □

With Lemmas 4, 5 and 6, we can conclude the following theorem.

THEOREM 5: For any set V of n nodes on $\Re$, for all $0 \le r \le 1$ and $\alpha \ge 2$,

$$\rho(NG_r(V)) \le 1 + r^{\alpha}(n-2) = F(r) .$$

Although this bound is related to the node size n so that $\rho\,(NG_r(V))$ can not be bounded, it can still be constant when r is 0 or some sufficiently small. i.e. $\rho\,(NG_r(V))$ is bounded in some range of r. So, we can make the following conclusion.

COROLLARY 1: The power stretch factor of the r-neighborhood graph is partially bounded.

Consider the maximum node degree of the r-neighborhood graph. Since $NG_r(V)$ consists of all edges in $RNG(V)$, the maximum node degree of $NG_r(V)$ is no less than that of $RNG(V)$. In section III, we know that $d_{\max}(RNG(V))$ is not always bounded in any case of V. Thus, $d_{\max}(NG_r(V))$ is also unbounded. Fortunately, Theorem 1 indicates that $d_{\max}(RNG(V))$ is bounded in most cases of V, where AS is assumed. Therefore, in the following theorem, we analyze the maximum node degree of the r-neighborhood graph under assumption AS.

THEOREM 6: For any set V of nodes on $\Re^2$ with assumption AS, for all $0 \le r \le 1$,

$$d_{\max}(NG_r(V)) \le \left\lfloor \pi / \sin^{-1}(r/2) \right\rfloor.$$

Proof. To prove this statement, it is sufficient to show that in $NG_r(V)$, there are no adjacent edges enclosing an angle less than $2\sin^{-1}(r/2)$. Assume for contradiction that two edges uv and uw in $NG_r(V)$ enclose an angle $\theta < 2\sin^{-1}(r/2)$ at node u, where $w, v \in V$. Without a loss of generality, we assume that $||uw|| < ||uv||$. With assumption AS, all nodes are placed on different positions. i.e. $P(x) \ne P(y)$, for any two nodes $x, y \in V$;

Consider the length of vw: If $\angle uwv$ is obtuse, it is clear that $||vw|| < ||uv||$ (note that $||vw||$ can not be equal to $||uv||$, since $P(u) \ne P(w)$), see Fig. 4(b); Otherwise, if $\angle uwv$ is not obtuse, $||vw||$ is less $||vw'||$, where $||uw'|| = ||uv||$, see Fig. 4(a). By the law of cosine, we have

$$\begin{aligned} \|vw'\|^2 &= \|uw'\|^2 + \|uv\|^2 - 2\|uw'\|\|uv\|\cos\theta \\ &= 2\|uv\|^2 - 2\|uv\|^2\cos\theta \\ &< 2\|uv\|^2 - 2\|uv\|^2\cos\left(2\sin^{-1}(r/2)\right) \end{aligned} \qquad \text{Eq. 2}$$

Let $\theta' = 2\sin^{-1}(r/2)$, we get $\sin(\theta'/2) = r/2$. Then one of the corresponding right-angled triangles is as shown in Fig. 4(c). In this case, $\cos\theta' = (2 - r^2)/2$. Thus we can get that $2\sin^{-1}(r/2) = \theta' = \cos^{-}\left((2-r^2)/2\right)$. Consequently,

$$\begin{aligned} \text{Eq. } 2 &= 2\|uv\|^2 - 2\|uv\|^2\cos\left(\cos^{-}\left(\left(2-r^2\right)/2\right)\right) \\ &= 2\|uv\|^2 - 2\|uv\|^2\left(\left(2-r^2\right)/2\right) = \|uv\|^2 r^2 \end{aligned} \qquad \text{Eq. 3}$$

Consequently, we have that for any case of $\angle uwv$,

$$\|vw\| < \max\{\|uv\|r, \|uv\|\} = \|uv\| \qquad \text{Eq. 4}$$

Consider the length of um: if $\angle uwm$ is obtuse, $||wm|| < ||uv||/2$ see Fig. 4(b); Otherwise, $||mw||$ is less $||mw'||$, see Fig. 4(b). By the law of cosine, we have

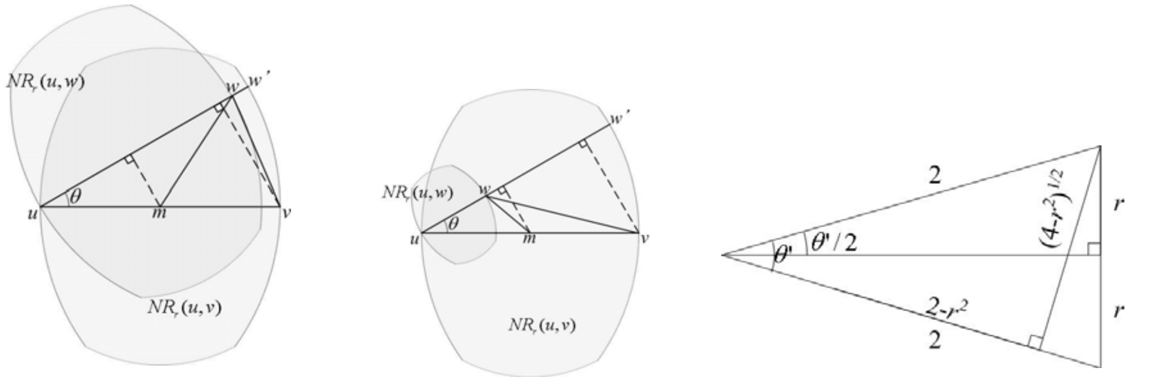
$$\begin{aligned} \|mw'\|^2 &= \|uw'\|^2 + \|um'\|^2 - 2\|uw'\|\|um'\|\cos\theta. \\ &< \|uv\|^2 + \|uv\|^2/4 - \|uv\|^2\cos\theta \\ &< 5\|uv\|^2/4 - \|uv\|^2\left(\left(2-r^2\right)/2\right) = \|uv\|^2\left(\left(1+2r^2\right)/4\right) \end{aligned} \qquad \text{Eq. 5}$$

Similarly, we have for any case of $\angle uwm$

$$\|mw\| < \max\left\{\|uv\|\sqrt{1+2r^2}/2, \|uv\|/2\right\} = l. \qquad \text{Eq. 6}$$

By Eq. 4, Eq. 6 and the assumption of $||uw|| < ||uv||$, w is included in the set of points specified in Eq. 1. Therefore, $P(w) \in NR_r(u, v)$. It however contradicts the assumption that uv is in

$NG_r(V)$. Thus we conclude this theorem. □



(a) (b) (c)

Fig. 4: (a) ∠*uwv* and ∠*uwm* are not obtuse; (b) ∠*uwv* and ∠*uwm* are obtuse; (c) a right-angled triangle with angle $\theta = 2\sin^{-}(r/2)$.

However, for those instances of *V* without *AS*, Theorem 6 can not hold anymore. See the instance in Fig 5, all nodes except v_i are placed on the outlier of $NR_r(v_i, v_1)$. This will result $n - 1$ neighbors adjacent to v_i in $NG_r(V)$

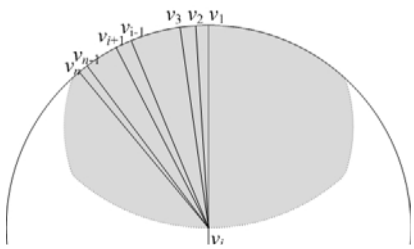


Fig. 5: $d_{\max}(NG_r(V))$ is not bounded if assumption *AS* does not hold.

So, in the next section, we propose an extended version the *r*-neighborhood graph. As the readers will see, the extended structure has the partially bounded maximum node degree for all cases of *V* and inherits almost all desired features in $NG_r(V)$.

IV. The Extended *r* –Neighborhood Graph

In this section, an extended structure of the *r*-neighborhood graph is given. The main goal is to avoid the unbounded maximum node degree in $NG_r(V)$. In this extension, assumption *AS* is not required anymore. Instead, a unique identifier $id(u)$ is available to each node *u* in *V*. The structure is defined as follows.

DEFINITION 4: Given a set *V* of nodes $\Re^2$, the *extended r-neighborhood graph* of *V*, denoted as $NG_r^*(V)$, has an edge *uv* if and only if $\|uv\| \le 1$ and there exists no node $w \in V$ satisfying one of the following three conditions:

D$_1$: $P(w) \in NR_r(u,v)$;

D$_2$: $P(w) \in D(m_{uv}, l_{uv}) \cap C(v, \|uv\|)$ and $id(u) > id(w)$;

D$_3$: $P(w) \in D(m_{uv}, l_{uv}) \cap C(u, \|uv\|)$ and $id(v) > id(w)$.

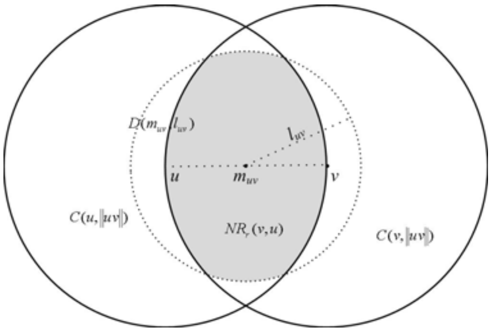


Fig. 6: The *r*-neighbor region of nodes *u* and *v*, and the two intersections defined in D$_2$ and D$_3$.

Without D$_2$ and D$_3$, $NG_r^*(V)$ is clearly equivalent to the original *r*-neighborhood graph. In conditions D$_2$ and D$_3$, the two sub-regions of $D(m_{uv}, l_{uv})$ intersected by $C(v, \|uv\|)$ and $C(u, \|uv\|)$ are, as depicted in Fig. 6 , the solid left arc and right arc along the outlier of $NR_r(u, v)$, respectively. When a node *w* is located in these two arcs, the existence of edge *uv* should be further determined by their identifiers.

Hereafter, we say that a node $w \in V$ *blocks* an edge *uv* in *UDG*(*V*) if and only if *w* satisfies one of the three conditions in Definition 4.

In $NG_r^*(V)$, an edge *uv* of *UDG*(*V*) will not only be blocked by some node *w* in $NR_r(u,v)$, but may also be blocked when either D$_2$ or D$_3$ happens. Therefore, $NG_r^*(V)$ constitutes a subgraph of $NG_r(V)$, which means that the maximum node degree of $NG_r^*(V)$ is no worse than its original version. In the following theorem, we show that the upper bound of $d_{\max}(NG_r(V))$ in Theorem 6 remains correct in $d_{\max}(NG_r^*(V))$, and the correctness is for any case of *V*, not subject to assumption *AS*.

THEOREM 7: For any set *V* of nodes on $\Re^2$, for all $0 \le r \le 1$,

$$d_{\max}(NG_r^*(V)) \le \left\lceil \frac{\pi}{\sin^{-1}(r/2)} \right\rceil.$$

Proof. Using the same argument as Theorem 6, we assume for contradiction that two edges *uv* and *uw* in $NG_r^*(V)$ enclose an angle $\theta' < 2\sin^{-1}(r/2)$ at node *u*. Without loss of generality, we assume that $\|uw\| \le \|uv\|$. If $\|uw\| < \|uv\|$, the argument of Theorem 6 has proved the contradiction. Consider $\|uw\| = \|uv\|$: Let w' be a point crossed by $C(u, \|uv\|)$ and the outlier of $D(m_{uv}, l_{uv})$, as shown in Fig. 7. The two edges $w'u$ and uv enclose an angle θ'. By the law of cosine, we have

$$\cos\theta' = \frac{\|uw'\|^2 + (\|uv\|/2)^2 - \|m_{uv}w'\|^2}{\|uw'\|\|uv\|} = \frac{\|uv\|^2 + (\|uv\|/2)^2 - l_{uv}^{2\ 2}}{\|uv\|\|uv\|} = 1 + r^2/2$$

Then one corresponding right-angle triangulation is as Fig.4(c). In this case, $\sin(\theta'/2) = r/2$. Thus, we can get that $\theta < \theta' = 2\sin^{-1}(r/2)$. Since $\|uw\| = \|uv\|$, both $P(w)$ and $P(v)$ are on $C(u, \|uv\|)$. The fact that $\theta < \theta'$ further limits $P(w)$ on the arc intersected by $D(m_{uv}, l_{uv})$. Similarly, $P(v)$ is limited on the arc intersected by $D(m_{uw}, l_{uw})$ for the same reason. Therefore, $P(w)$ and $P(v)$ are on the regions defined in D_2, with respect to edges uw and uv, respectively.

Next, the existence of uv and uw should be determined by their identifiers. If $id(v) > id(w)$, uv is blocked by w. Otherwise, if $id(v) < id(w)$, uw is blocked by v. As a sequel, no matter what the values of $id(v)$ and $id(w)$ are, at least one of the edges enclosing θ' can not be in $NG_r^*(V)$. Thus we proved this theorem. □

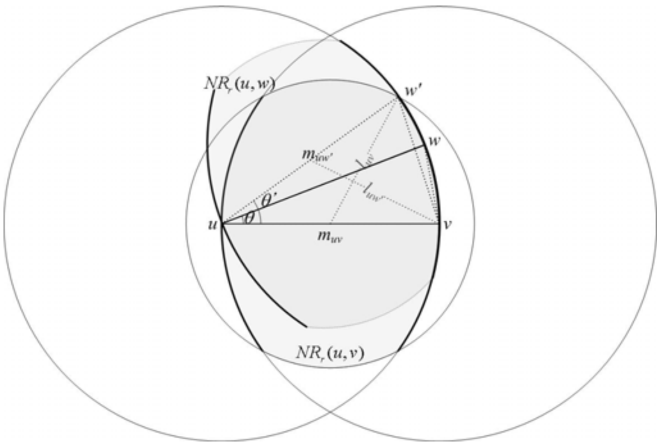


Fig. 7: If $\theta < 2\sin^{-1}(r/2)$ and $\|uw\| = \|uv\|$, either uw or uv can not be in $NG_r^*(V)$

From Theorem 7, we can see that $d_{\max}(NG_r^*(V))$ is constant when r is sufficiently large. Therefore, there has some setting of r such that $d_{\max}(NG_r^*(V))$ is bounded by some constant, for any set V of n nodes. So, we reach the following conclusion.

COROLLARY 2: The maximum node degree of the extended r-neighborhood graph is partially bounded.

In the rest part, we show that $NG_r^*(V)$ inherits all desired properties achieved by

$NG_r(V)$, except the generality for $RNG(V)$. The fact that $NG_r^*(V) \subseteq NG_r(V)$ confirms the planarity of $NG_r^*(V)$, since $NG_r(V)$ is planar for any r. Moreover, when $r = 0$, the two arcs defined in D_2 and D_3 are empty. Thus whether an edge is in $NG_r^*(V)$ is solely depending on D_1, which means that $NG_0^*(V) \equiv NG_0(V) \equiv GG(V)$. Therefore, $NG_r^*(V)$ remains a general structure of $GG(V)$.

However, as shown Theorem 7, some adjacent edges having the same length in $RNG(V)$ would be avoided in $NG_r^*(V)$. Thus $RNG(V)$ is not always a subgraph of $NG_r^*(V)$. This means that $NG_1^*(V)$ is not essentially equivalent to $RNG(V)$. Even more, $NG_1^*(V)$ could be a subgraph of $RNG(V)$. Therefore, $NG_r^*(V)$ is no longer a general structure of $RNG(V)$.

About the connectivity, because $RNG(V)$ is not always a subgraph of $NG_r^*(V)$, we cannot ensure the connectivity of $NG_r^*(V)$ directly from that of $RNG(V)$. Therefore, we apply an entirely different logic to prove this property. The idea is based on comparing the lexicography orders of nodes pairs. This idea has been successfully used to prove the connectivity of $XTC(V)$ [32], another subgraph of $RNG(V)$.

We define a three-field tuple ($\|uv\|$, $id(u)$, $id(v)$) for each nodes pair (u, v). The lexicographic order of (u, v) is smaller than that of another nodes pair (s, t) if one of the following three cases happens: (1) $\|uv\| < \|st\|$; (2) $\|uv\| = \|st\|$ and $id(u) < id(s)$; (3) $\|uv\| = \|st\|$, $id(u) = \mathrm{id}(s)$ and $id(v) < id(t)$. Now, we prove the connectivity of $NG_r^*(V)$ in Theorem 8.

THEOREM 8: For any set V of nodes on $\Re^2$, if the underlying $UDG(V)$ is connected, $NG_r^*(V)$ is *connected*, for all $0 \le r \le 1$.

Proof. Suppose $UDG(V)$ is connected. Let $U(V)$ be the set of unconnected nodes pairs in $NG_r^*(V)$. We assume for contradiction that some nodes pairs in $NG_r^*(V)$ are not connected. i.e., $U(V)$ is not empty. Let (u, v) be the node pair with smallest lexicographic order in $U(V)$. Assume that edge uv is not in $UDG(V)$, i.e. $\|uv\| > 1$. Since $UDG(V)$ is connected, there must be some path longer than one hop connecting u and v. Let $\pi(u, v)$ be such path in $UDG(V)$. Since $\|uv\| > 1$, the lengths of each edge on $\pi(u, v)$ is less than $\|uv\|$. When this path is mapped to $NG_r^*(V)$, there is some nodes pairs on $\pi(u, v)$ unconnected in $NG_r^*(V)$. Thus some unconnected node pair on $\pi(u, v)$ has length shorter than $\|uv\|$, which however contradicts that (u, v) has the smallest lexicographic order in $U(V)$. Therefore, edge uv must be in $UDG(V)$.

Since edge uv is in $UDG(V)$ and not in $NG_r^*(V)$, there must be some node w satisfying one of the three conditions in Definition 4. Besides, either (u, w) or (w, v) is in $U(V)$, otherwise (u, v) can be connected by path uwv. We consider the three cases:

1) If D_1 happens, $P(w) \in NR_r(u, v)$. So, we has $||uw|| < ||uv||$ and $||wv|| < ||uv||$, which means that the lexicographic orders of (u, w) and (w, v) are less than that of (u, v).
2) If D_2 happens, we have $||wv|| = ||uv||$ and $id(u) > id(w)$, which means that the lexicographic order of (w, v) is less than that of (u, v);
3) If D_3 happens, we have $||uw|| = ||uv||$ and $id(v) > id(w)$, which means that the lexicographic order of (u, w) is less than that of (u, v).

Therefore, we cannot find any nodes pair in $U(V)$ having the smallest lexicographic order. In other words, $U(V)$ is empty, which however is a contradiction. Thus, we proved this. □

Due to the fact that $NG_r^*(V) \subseteq NG_r(V)$, there may has some paths in $NG_r(V)$ not in $NG_r^*(V)$. Therefore, $\rho(NG_r^*(V))$ is no better or even worse than $\rho(NG_r(V))$. Even so, the upper bound of $\rho_{NG_r^*(V)}(UDG(V))$ can be as good as that proved in Theorem 5; we briefly explain this: All arguments in Theorem 5 are not related to the two additional conditions D_2 and D_3, except those referred from Lemma 2. Whatever D_1, D_2 or D_3 happens, $||uw|| \le ||uv||$, $||vw|| = ||uv||$ and $||mv|| < l$, which means that all inequalities in the proof of Lemma 2 are unchanged. Consequently, Theorem 5 is still correct, even if all conditions of Definition 4 are considered. So, $\rho(NG_r^*(V))$ is also partially bounded.

Below, we show that the bound $1 + r^\alpha(n-2)$ in Theorem 5 is not only correct, but also *asymptotically tight* to the worst possible value of $\rho_{NG_r^*(V)}(UDG(V))$. In other words, it is very hard to find another upper bound of $\rho_{NG_r^*(V)}(UDG(V))$ better than ours. We apply the same argument as that used to verify the tightness of the length stretch factor [3] and the power stretch factor [15] of $RNG(V)$

THEOREM 9: For any $n \ge 2$ and $0 \le r \le 1$, there is a set V of n nodes such that

$$\sup_{|V|=n} \rho_{NG_r^*(V)}(UDG(V)) > 1 + r^\alpha(n-2) - \varepsilon,$$

for any sufficient small $\varepsilon > 0$.

Proof. Let $\theta_1 = 2\sin^{-1}(r/2) - 2\lambda$ and $\theta_2 = \pi/2 - \sin^{-1}(r/2) + \lambda$, where $\lambda > 0$. We construct a set

$V = \{v_1, v_2, \ldots, v_{2m-1}, v_{2m}, \ldots, v_n\}$ of n nodes, where $n \geq 2$ is even and $m = n/2$, as follows:

1) $\|v_1v_2\| \leq 1$ and $\|v_iv_{i+1}\| = \|v_1v_2\|$, for $i = 2, 3, \ldots, 2m-1$;
2) $\angle v_iv_{i+1}v_{i+2} = \theta_1$, for $i = 1, 2, \ldots, 2m-2$;
3) $\angle v_{i+2}v_iv_{i+1} = \angle v_iv_{i+2}v_{i+1} = \theta_2$, for $i = 1, 2, \ldots, 2m-2$;
4) $id(v_i) = n - i + 1$, for $i = 1, 2, \ldots, n$;

One corresponding $UDG(V)$ is as shown in Fig. 8 (a). For $i = 1, 2, \ldots, 2m-2$, since $\angle v_iv_{i+1}v_{i+2} = \theta_1 < 2\sin^{-1}(r/2)$ and $\|v_iv_{i+1}\| = \|v_{i+1}v_{i+2}\|$, by the argument in Theorem 7, we get $P(v_{i+2}) \in D(v_i, v_{i+1}) \cap C(m_{v_iv_{i+1}}, l_{v_iv_{i+1}})$. That is, $P(v_{i+2})$ is in the regions with respect to edge v_iv_{i+1}, defined in D_2. Moreover, $id(v_i) > id(v_{i+2})$. Thus, edge v_iv_{i+1} is not in $NG_r^*(V)$. Then, the remaining edges are exactly a path (spanning tree) $v_1v_3v_5\ldots v_{2m-3}v_{2m-1}v_{2m}v_{2m-2}\ldots v_6v_4v_2$ of V, connecting all nodes, as the bold links in Fig. 8 (a). Therefore, we can get that

$$p\left(\pi^*_{NG_r^*(V)}(v_1, v_2)\right) = \sum_{i=1}^{2m-2} r^\alpha \|v_iv_{i+2}\|^\alpha + \|v_{2m-1}v_{2m}\|^\alpha$$

As $\lambda \to 0$, $\theta_1 \to 2\sin^{-1}(r/2)$, which implies that $\|v_iv_{i+2}\| \to r\|v_iv_{i+2}\| = r\|v_1v_2\|$, according to Eq. 3. Consequently, as $\lambda \to 0$, we get that

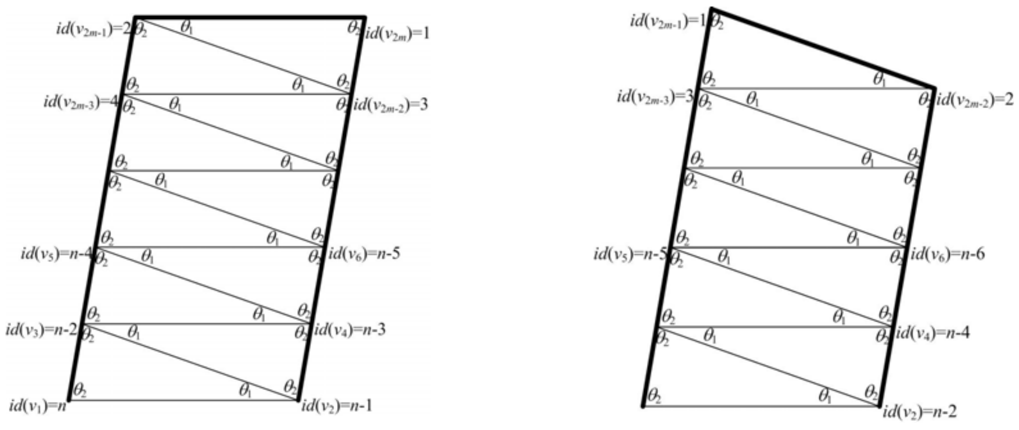
$$\begin{aligned} &\sum_{i=1}^{2m-2} r^\alpha \|v_iv_{i+2}\|^\alpha + \|v_{2m-1}v_{2m}\|^\alpha \\ &\to \sum_{i=2}^{2h-2} r^\alpha \|v_1v_2\|^\alpha + \|v_1v_2\|^\alpha \\ &= \|v_1v_2\|^\alpha \left((n-2)r^\alpha + 1\right) \end{aligned}$$

On the other than, since $\|v_1v_2\| \leq 1$, we get $p\left(\pi^*_{UDG(V)}(u, v)\right) = \|uv\|^\alpha$. Therefore, as $\lambda \to 0$, $\rho_{NG_r(V)}(UDG(V)) \to 1 + r^\alpha(n-2)$. That is, $\sup_{|V|=n} \rho_{NG_r^*(V)}(UDG(V)) > 1 + r^\alpha(n-2) - \varepsilon$, for any sufficient $\varepsilon > 0$. For any odd $n \geq 2$, the result can be obtained by applying the same argument to the instance as shown in Fig. 8(b). So, we proved this theorem. □

Actually, an equivalent structure of $NG_r^*(V)$, without a original version like $NG_r(V)$, was mentioned in our previous paper[3] [9]. In that preliminary work, however only qualitative results were given. To prove the quantitative results, we separate $NG_r(V)$ from $NG_r^*(V)$ in this paper, because $NG_r(V)$ has a clearer form in definition that can be used to highlight the

[3] The term "r-neighborhood graph" in [9], is not refereed to the original version in Definition 3, but the extended version in Definition 4. In this paper, we reuse the same term to name the original version and rename the previous structure in [9] the extended version

main tricky in our derivations. Besides, all qualitative results in [9] are re-evaluated here using different arguments.



(a) (b)

Fig. 8: A worst-case instance V of n nodes in $NG_r^*(V)$: (a) n is even; (b) n is odd.

VI. Purely Localized Algorithm

In this section, we propose an efficient purely localized algorithm, named *PLA*, to construct the r-neighborhood graph. This algorithm consists of two main procedures, GETINF and FINDNB. First, GETINF collects a set of nodes' information within one-hop distance, denoted as IN_u. Then, the collected information will be fed into FINDNB to determine a set of neighbors in $NG_r(V)$, denoted as NB_u.

ALGORITHM *PLA*
Input: A ratio $0 \le r \le 1$.
Output: A set of neighbors adjacent to u.
Step 1: IN_u := GETINF(u, r);
Step 2: NB_u := FINDNB(u, r, IN_u);
Step 3: Stop and output NB_u;

To collect the one-hop information, the simplest way is to let each node broadcast its information at the maximum transmission range 1 and gather the information from others. However, the severe path loss and the frequent change in topology may cause considerable power in such transmission. Therefore, in GETINF we aim to reduce the transmission range during construction. The main idea is to incrementally raise the transmission power from a small range and then use some rule to stop the increment earlier before the transmission range 1 is reached. The detail steps are explained as follows: the transmission range is initiated at a small distance d_0, and then it will be incrementally raised for several rounds. Let d_1 and d_2 be the previous and the current transmission ranges of a round respectively. In each round, a node broadcasts a request to distance d_2, and waits for the responses from receiving nodes to gather the nodes' information. To avoid replying to a node for the second times, the request of a node u contains the position $P(u)$ and the previous distance d_1. As a node v receives this request, it calculates the Euclidean distance $\|uv\|$. Then, if $\|uv\| > d_1$, v responses its information, $P(v)$, to u at distance $\|uv\|$, otherwise, just neglects the request. In each round, the range is increased by multiplying $\sqrt[\alpha]{2}$, which means the transmission power is multiplied by 2 each time. The process is continued until the following stopping criterion is satisfied. Let v_1 and v_2 be two crossed points intersected by $C(u, \|uv\|)$ and $C(m, l)$, see Fig. 9 (a). We define $SC(u, v)$ to be the semicircle enclosed by uv_1 and uv_2 with radius ε, where $\varepsilon > 0$ is a small value less than the distance between any pair of nodes in V. Then, given a distance d, a semicircle $\chi(u, d)$ is defined as follows

$$\chi(u,d) = \bigcup_{\|uv\| \le d} SC(u,v) .$$

We can prove that if $\chi(u, d)$ is exactly the circle $C(u, \varepsilon)$, like Fig. 9(b), then a disk centered at u with d radius can cover all neighbors of u in $NG_r(V)$. In other words, GETINF can be halted as $\chi(u, d_2) \equiv C(u, \varepsilon)$. Let $N_u(G(V))$ be the set of neighbors of node u in a graph $G(V)$. This property is proven in Lemma 7.

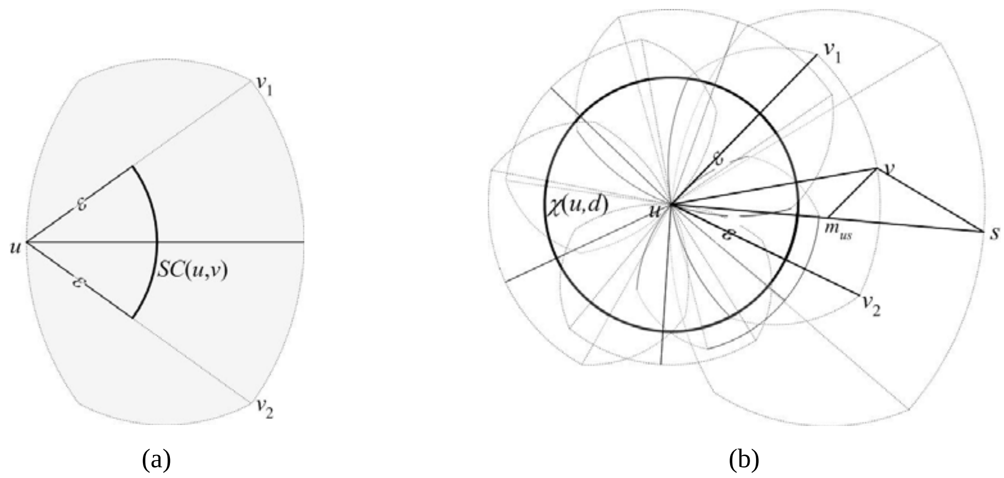


Fig. 9: (a) the semicircle $SC(u, v)$; (b) the $\chi(u, d)$ is the union of all $SC(u, v)$ where v is within distance d.

LEMMA 7: Given a node $u \in V$ and distance $d \in \Re$, if $\chi(u,d) \equiv C(u,\varepsilon)$,

$$N_u(NG_r(V)) \subseteq \{v \in V \mid P(v) \in D(u,d)\}.$$

Proof. We assume for contradiction that some node s in $N_u(NG_r(V))$ is not in $\{v \in V \mid P(v) \in D(u,d)\}$. Since ε is less than the distance between any pair of nodes in V, we get $||us|| > \varepsilon$. Thus, edge us intersects a point on the circle $C(u, \varepsilon)$. Do the fact that $\chi(u,d) \equiv C(u,\varepsilon)$, us must intersect at least one semicircle that composes $\chi(u, d)$, see Fig. 9 (b). Let $SC(u, v)$ be one of the semicircles intersected by us. Then, us is enclosed by uv_1 and uv_2 in $SC(u, v)$. In other words, $\angle suv \le \angle v_1uv$ or $\angle suv \le \angle vuv_2$. According to the argument in Theorem 7, we can get that $\angle v_1uv = \angle vuv_2 = 2\sin^{-1}(r/2)$. Therefore, we have $\angle suv \le 2\sin^{-1}(r/2)$. Moreover, since s is not in $\{v \in V \mid P(v) \in D(u,d)\}$, s must be farer than v from u. So, $P(v) \in NR_r(u, s)$. According to Definition 4, us in not in $NG_r(V)$, which however contradicts that s is a neighbor of u in $NG_r(V)$. Thus, we concluded this lemma. □

The total transmission power used by GETINF could be as large as $d_0^\alpha(1+2^1+2^2+\cdots+2^I)$, where I is the number of rounds. This result could be worse than the maximum transmission power 1 as I is large. Fortunately, when n is large, nodes are closer to and evenly surrounded by each other so that χ (u, d) has more change to be quickly shaped as $C(u, \varepsilon)$. So we can gain benefit from GETINF in higher probability as the number of nodes increases.

The steps of GETINF are described below. Neglecting the communication overhead at step 2, the execution time of GETINF is dominated by the union operation at step 4. This step can be implemented by some search-and-merge algorithm. Thus, the time complexity of GETINF is O($n$log$n$).

GETINF(u, r)
Step 1: $d_1 := 0$, $d_2 := d_0$, $IN: = \phi$, $\chi(u, d_2): = \phi$;
Step 2: Broadcast a request ($P(u)$, d_1) to distance d_2 and gather a set R of responses from nodes within d_1 and d_2;
Step 3: For each $v \in R$ do
    $\chi(u,d_2) := \chi(u,d_2) \cup SC(u,v)$;
Step 4: $IN: = IN \cup R$;
Step 5: If $d_2 \le 1$ and $\chi(u, d_2)$ is not the circle $C(u, \varepsilon)$ do
    $d_1 = d_2$;
    $d_2 := d_2 \times 2^{1/\alpha}$;
    Return to step 2;
Step 6: Stop and output IN;

Now we discuss the communication cost of GETINF. As d_0 is multiplied by $\sqrt[\alpha]{2}$ over

$\alpha\log_2(1/d_0)$ times, it is larger than 1. Therefore, the number of rounds to increase the transmission range d_2 is dominated by $\alpha\log_2(1/d_0) + 1$. Assume a node's position can be encoded by $\log_2 n$ bits. Each node has to broadcast at most $(\log_2 n)(\alpha\log_2(1/d_0) + 1)$ bits for the request messages. In addition, a node will reply to same node no more than once. Thus, a node needs at most $(\log_2 n)(n - 1)$ bit to reply all requests. Combining these results, communication cost of a node is no more than $(\log_2 n)(\alpha\log_2(1/d_0) + n)$ bits.

Once the information IN_u is collected, node u can start to determine its neighbors in $NG_r(V)$. One institutive way is to apply Definition 2 on IN_u directly, as the follows procedure.

Step 1: $N := IN_u$;
Step 2: For each node v in N do
        For each node $w \in IN_u$ do
            If $P(w) \in NR_r(u, v)$ do
                $N := N - \{v\}$;
Step 3: Output N and stop;

In this procedure, the existence of a neighbor v in IN_u is determined by checking whether some node w is located in $NR_r(u, v)$. The correctness is obvious, while in the worst case it should take $O(n^2)$ time on each node. This time is usually not tolerable when topology changes frequently. Therefore, we aim to reduce time complexity in this part. In FINDNB, the main idea is to reverse the original procedure. That is, instead of checking whether some node w can block an edge uv, for each uv, we check whether some edge uv can be blocked by a node w, for each w. The procedure is below.

This checking is begun from the farthest to the closet nodes in IN_u. So, we index all elements of IN_u in non-decreasing order of $\|uw\|$ in step 2. The set NB contains all candidates that could be a neighbor of u during the process. As a node w is given, we remove from NB all fail candidates that that are already blocked by w. After that, w is added into NB to be an new candidate of $N_u(NG_r(V))$. The process continues until all w's in IN_u were considered. Now, we prove the correctness of FINDNB.

FINDNB(u, r, IN_u)
Step 1: $NB := \phi$;
Step 2: Index the elements of IN_u in non-increasing order of $\|uw\|$;
Step 3: For each node $w \in IN_u$ with *smallest* index do

For each node $v \in NB$ do
If $P(w) \in NR_r(u,v)$ do
$NB := NB - \{v\}$;
$NB := NB + \{w\}$;
Step 4: Stop and output NB;

THEOREM 10: For any set V of nodes on $\Re^2$, $NB_u = N_u(NG_r(V))$, for any $u \in V$.

Proof. We prove this by showing that for any $v \in V$, $v \in NB_u$ if and only if edge uv is in $NG_r(V)$. Suppose an edge uv is in $NG_r(V)$. By Definition 2, there is no $w \in N_u(UDG(V))$ such that $P(w) \in NR_r(u, v)$. This implies that once v is added in NB, there is also no $w \in IN_u$ such that v can be removed at step 3. Since $v \in N_u(NG_r(V)) \subseteq IN_u$ and each node in IN_u can be added to NB, v must be in NB at least one time. So, we can get that v is in the final output of NB_u. Contrarily, we suppose $uv \notin NG_r(V)$. Some node $w \in N_u(UDG(V))$ is located in $NR_r(u, v)$. If $v \notin IN_u$, the result clearly follows by Lemma 7. Otherwise, $v \in IN_u$. In this case, all nodes blocking uv are in IN_u. Besides, every node w blocking uv is always considered after v in GETNB. Therefore, even if v can be added to NB, there must be a node $w \in IN_u$ such that v can be removed from NB at the successive iteration. So we get $v \notin IN_u$. □

Lemma 7 also implies that if $uv \in NG_r(V)$, then $v \in N_u$ and $u \in N_v$ and that if $uv \in NG_r(V)$, then $v \notin N_u$ and $u \notin N_v$. So, the neighbors (links) determined by GETNB are symmetric.

COROLLARY 3: Any topology resulted by *PLA* is symmetric.

Consider the time complexity of FINDNB. Step 2 can be done by some sorting algorithm in $O(n\log n)$. Before a node $w \in IN_u$ is added to NB, any $v \in IN$ blocked by w is removed from NB. Therefore, for any two nodes in NB, none of them can be blocked by each other. Let s and t be two nodes in NB. The argument of Theorems 7 indicates that if $\angle sut < 2\sin^{-1}(r/2)$, then either s blocks t or t block s. Since neither s blocks t nor t blocks s, we get that $\angle sut \geq 2\sin^{-1}(r/2)$. Therefore, during the process, the size of NB can be never greater than $d_{\max}(NG_r(V))$. Consequently, FINDNB can be done in $O(n\max\{\log n, d_{\max}(NG_r)\})$ time. We can observe that this time complexity depends on the parameter r. When r equals or closes to 0 (the worst cases), the time complexity of FINDNB is still $O(n^2)$. However, when r is sufficiently large such that $d_{\max}(NG_r(V))$ is a constant, FINDNB can be done in $O(n\log n)$.

With a slight modification, *PLA* can be easily applied on the extended r-neighbors graph and all results can be preserved. We omit the detail explanation here.

VII. Conclusion

In this paper, we proposed a purely localized structure to control the topology in wireless networks. We showed the worst case of the power stretch factor is an increasing function of r and the worst cast of the maximum node degree is contrarily a decreasing function of r. So, the two objectives can be adjusted in our structure. Although the power stretch factor is related to n so that our structure is not really a spanner, $\rho(NG_r(V))$ can still be bounded for some range of r. Therefore, the power stretch is partially bounded in our structure. About the maximum node degree, we proposed an upper bound derived for $d_{\max}(NG_r(V))$. However, this result is correct only no node having two or more neighbors at exactly distance. For this reason, an extended structure $NG_r^*(V)$ was given to comprehend this theorem.

Besides, the proposed structure can always result connected topology with symmetric edges. Any resulting topology is always a planar. The relations between the r-neighborhood graph and existent structures are summarized as follows. Specially, $NGr(V)$ is a general structure of both $GG(V)$ and $RNG(V)$.

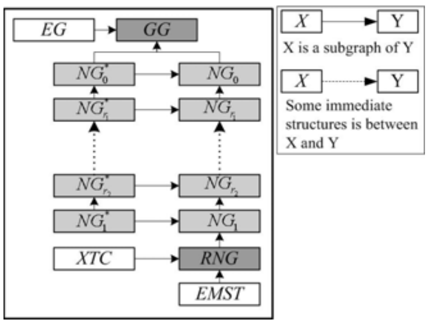


Fig. 10 The relationships of $NG_r(V)$, $NG_r^*(V)$, $GG(V)$ and $RNG(V)$.

To construct our structure, we proposed a 1-hop purely localized algorithm, *PLA*. It can avoid long-distance transmission when collecting information and can be efficiently done in O($n$log$n$) time when $d_{\max}(NG_r(V))$ is constant.

For the further research, a localized topology control approach enables the design of localized routing protocols. For instance, the greedy route discovery in CFG [26] and GPSR [11] are based on *GG*. We anticipate that r-neighborhood graph could provide a concrete basis for many interesting extensions due to the sound theoretical results. Moreover, the parameter r can be turned to find the best settings for different scenarios. Another interesting issue for the possible further work is to evaluate the stability of the proposed structure when perfect position (range) information is not available or when the accuracy of position information differs from node to nodes.

Appendix

The proof of Lemma 2: Without a loss of generality, we assume that ||*uw*|| ≤ ||*vw*||. Let *y* be the projection of *w* on *uv* so that *yw* is perpendicular to *uv*. We can derive that

$$\|ym\| = \frac{\|vw\|^2}{2\|mv\|} - \frac{\|wm\|^2}{2\|mv\|} - \frac{\|mv\|}{2} \quad \text{and} \quad \|yx\| = \sqrt{\|wv\|^2 - \left(\frac{\|vw\|^2}{2\|mv\|} - \frac{\|wm\|^2}{2\|mv\|} + \frac{\|mv\|}{2}\right)^2}\ .$$

Thus,

$$\begin{aligned}
\|uw\|^2 &= \|wy\|^2 + (\|um\| - \|ym\|)^2 = \|wy\|^2 + (\|mv\| - \|ym\|)^2 \\
&= \|vw\|^2 - \left(\frac{\|wv\|^2}{2\|mv\|} - \frac{\|wm\|^2}{2\|mv\|} + \frac{\|mv\|}{2}\right)^2 + \left(\frac{3\|mv\|}{2} - \frac{\|vw\|^2}{2\|mv\|} + \frac{\|wm\|^2}{2\|mv\|}\right)^2 \\
&= 2\|mv\|^2 - \|vw\|^2 + 2\|wm\|^2
\end{aligned}$$

Then, power consumed by path *uwv* is as follows

$$p(uwv) = \|uw\|^\alpha + \|vw\|^\alpha = \left(2\|mv\|^2 - \|vw\|^2 + 2\|wm\|^2\right)^{\frac{\alpha}{2}} + \|vw\|^\alpha$$

From Eq.1 we get $\|wm\| < l = \|uv\|\sqrt{1+2r^2}\,/\,2 = \|mv\|\sqrt{1+2r^2}$ and $\|vw\| < \|uv\|$, so

$$\begin{aligned}
&\left(2\|mv\|^2 - \|vw\|^2 + 2\|wm\|^2\right)^{\frac{\alpha}{2}} + \|vw\|^\alpha \\
&\le \left((4+4r^2)\|mv\|^2 - \|vw\|^2\right)^{\frac{\alpha}{2}} + \|vw\|^\alpha \\
&\le \left(\frac{(4+4r^2)}{4}\|uv\|^2 - \|uv\|^2\right)^{\frac{\alpha}{2}} + \|uv\| \\
&= \left(r^2\|uv\|^2\right)^{\frac{\alpha}{2}} + \|uv\|^\alpha = \|uv\|^\alpha(1+r^\alpha)
\end{aligned}$$

Thus, we have that $p(uwv) \le \|uv\|^\alpha(1+r^\alpha)$ □

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An Kai Jeng received the B. S in Statistics from Tamkang university and the M.S. in Management Information System from National Chi Nan University in 2001 and 2003, respectively. He is currently pursuing the Ph.D. degree in the department of Computer and Information Science at National Chiao-Tung University, Taiwan, Republic of China. His research interests include wireless networks, algorithm design and analysis, scheduling theory and operation research.

Rong-Hong Jan received the B.S. and M.S. degrees in Industrial Engineering, and the Ph.D. degree in Computer Science from National Tsing Hua University, Taiwan, in 1979, 1983, and 1987, respectively. He joined the Department of Computer and Information Science, National Chiao Tung University, in 1987, where he is currently a Professor. During 1991-1992, he was a Visiting Associate Professor in the Department of Computer Science, University of Maryland, College

Park, MD. His research interests include wireless networks, mobile computing, distributed systems, network reliability, and operations research.

A Power-Saving Scheduling for Infrastructure-Mode 802.11 Wireless LANs*

Hsiao-Po Lin, Shih-Chang Huang and Rong-Hong Jan†

Department of Computer and Information Science, National Chiao Tung University, Hsinchu, 30050, Taiwan

Abstract

This paper presents a novel method to arrange wakeup schedule for sleeping stations such that the number of wakeup stations in each beacon interval is balanced in IEEE 802.11 wireless local area networks (WLANs). This method reduces the probability of collision and thus the station can save more power. Next, we consider how to poll the wakeup stations to send the PS-Poll frame to get back their buffered data so that the contention can be avoided. Three different access scheduling mechanisms are proposed for the contention avoidance. In the first mechanism, only one of wakeup stations is scheduled to access the buffered data. The second and third mechanisms based on the smallest association ID (AID) first and the smallest queue length first, respectively, arrange a subset of wakeup stations to get back their buffered data within a beacon interval. Simulation results show that the proposed methods are effective in the power-saving.

Keywords: Wireless LAN, Infrastructure mode, Power saving, Traffic scheduling.

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†Corresponding Author. Fax: 886-3-5721490; e-mail: rhjan@cis.nctu.edu.tw

1 Introduction

Recently, due to the technology explosion in wireless communication (e.g., Bluetooth, IEEE 802.11, GSM/GPRS, and WCDMA) and portable communication devices (e.g., notebook PCs, personal digital assistants, and smart phones), it has become possible for people to connect to the Internet anytime and anywhere, and remain on-line while roaming. Among these wireless communication techniques, IEEE 802.11 wireless local area networks (WLANs) [1, 2, 3] are the most widely-used local wireless network system in schools, offices, airports, etc. Besides, almost portable devices can be equipped to access IEEE 802.11 WLANs. However, the energy sources of these portable devices come from their equipped batteries. Once the battery has run down, it needs to be recharged and portable device suffers a loss of network connectivity. Thus, the power-saving problem becomes an important issue for prolonging the operation of a portable device [4] .

For IEEE 802.11 WLANs, a common method for power-saving is to powering down the transceiver. When transceiver is off, we say the mobile station is in *sleeping*. A listen interval is a period of time for which the mobile station may choose to sleep. Power conservation in IEEE 802.11 can be achieved by maximizing the listen interval. However, longer listen intervals increase the transmission delay. In [5] and [6], they proposed listen-interval adaptation mechanisms for power-saving in which the mobile station dynamically adapts the duration of listen-interval according to the traffic situation. Thus, if the traffic load is light, the mobile device can set a longer listen-interval to save more power. Another kind of listen interval controlling is the quorum based scheme [7, 8]. In the quorum based scheme, each station in power-saving mode synchronizes its wakeup schedule with each other such that the station can deliver buffered frames to a power-saving station at right time when the radio of power-saving station is turned on.

Controlling the transmission power is another way to save mobile devices power [6]. The basic idea is using the least power to transmit data. In [6], an adaptive transmission power mechanism is proposed in which the access point (AP) computes the optimal transmission power for each associated mobile station based on the received signal and informs the mobile devices their optimal transmission power. Then, the mobile station can adjust its transmission power to the optimal value to saving power.

Frame aggregation can also be used to save power. There are two kinds of frame aggregation. One is combining multiple relative frames into an aggregated frame [9]. The other is compressing the payload to reduce the amount of transmission data.

Another interesting way to saving power is using a low power radio module, such as bluetooth [10], or a special signaling channel [11], to actuate the normal IEEE 802.11 radio circuit. In [10], the bluetooth module is used for signaling and IEEE 802.11 for data transmission. This method indeed can save mobile station power but the cost is that it needs an extra hardware. The method in [11] is similar to that in [10] except the low power radio module.

TDMA-based (time-division multiple-access) approaches [12, 13] can also assist mobile stations to conserve power by scheduling channel access in advance so that mobile stations can turn their transceivers off when it is not their turn to transmit or receive. In [12], the access point periodically broadcasts a schedule frame containing start time and duration of time slot for each mobile station. The mobile station may send or receive frames only during its time slot and outside its time slots, the mobile station can turn its transceiver off to save power. In [13], they only focus on the point coordination function (PCF). When the PCF is used, time on the medium is divided into the contention free period (CFP) and the contention period (CP). In a CP, the point coordinator learns what traffic needs to be transmitted and then directs its transmissions in a CFP. In [14], they propose a power-saving algorithm that incorporates a contention-free scheduling function for data transmission in 802.11 ad hoc networks.

This paper focuses on the power management problem for an infrastructure mode IEEE 802.11 WLAN. We present a novel method to arrange wakeup schedule for sleeping stations such that the number of wakeup stations in each beacon interval is balanced in IEEE 802.11 WLANs. This method can reduce the probability of collision and thus the station saves more power. Next, we consider how to poll the wakeup stations to send the PS-Poll frame to get back their buffered data so that the contention can be avoided. Three different access scheduling mechanisms are proposed for the contention avoidance. In the first mechanism, only one of wakeup stations is scheduled to access the buffered data. The second and third mechanisms schedule a subset of wakeup stations to retrieve their buffered data within a beacon interval. The access sequences of the second and third mechanisms are based on the smallest association ID (AID) first and the smallest queue length first, respectively.

The rest of this paper is organized as follows. Section 2 reviews the operation of power saving mode in IEEE 802.11 WLANs. A wakeup scheduling problem is considered in section 3 and three contention avoidance mechanisms for polling wakeup stations are presented in section 4. In section 5, we give the simulation results to show the effectiveness of our proposed methods. Finally, the conclusion and possible future research are given in section 6.

2 Power Management in 802.11 WLANs

In infrastructure mode IEEE 802.11 WLANs, a mobile station can power down its transceiver and enter the power-saving mode (PS mode) for conserving power. The station can communicate its power management state to its AP. Thus, an AP knows the power management state of every station that has associated with it. When a frame arrives, the AP can determine whether the frame should be delivered to wireless network because the station is awake or buffered because the station is in PS mode.

After buffering frames, the next job for AP is to announce periodically which stations have frames waiting for them. That is, AP periodically broadcasts beacon frames with a traffic indication map (TIM) to its service stations. The TIM is a virtual bitmap in which each bit corresponds to a particular AID. When a station has associated to an AP, it receives an AID from the AP. The AP sets the bit in TIM if it has buffered frames for the station with AID corresponding to the bit position.

Mobile stations in power saving mode have to wake up to listen for beacon frames and check the TIM. By this way, a mobile station can determines whether the AP has buffered frames for it. If the AP seldom buffers frames for the station, the station does not require waking up to check every beacon frame. Instead, it wakes up every *listen interval* to check the beacon frame. A listen interval is a number of beacon interval for which the mobile station may choose to sleep. If the station finds that the AP has buffered data for it, it will send a PS-Poll control frame to retrieve the buffered frames. When multiple stations have buffered frames, all stations with buffered frame contend the medium for sending PS-Poll. After sending the PS-Poll, a station has to awake until the buffered frames are received or the bit in the TIM corresponding its AID is no longer set.

For example, as shown in Figure 1, a station, denoted as STA, is wakeup in the first beacon interval and receives the beacon frame in which the TIM indicates buffered data for it. Then, STA contends the medium for sending a PS-Poll frame to inform the AP that it is wakeup and ready to get back the buffered data. After AP receives the PS-Poll, it transmits a buffered frame to the STA. The STA returns an ACK frame to inform AP that the frame is received completely.

3 Load-Aware Wakeup Scheduling

Consider a wireless LAN having an AP and six sleeping stations, A, B, C, D, E, and F. The listen intervals of stations, A, B, C, D, E, and F are 1, 2, 3, 6, 6, and 6, respectively. Let $w_i(t)$ be an indication bit where $w_i(t) = 1$ if the station i wakes up at beacon interval t; $w_i(t) = 0$, otherwise. Let $n(t)$ denote the total

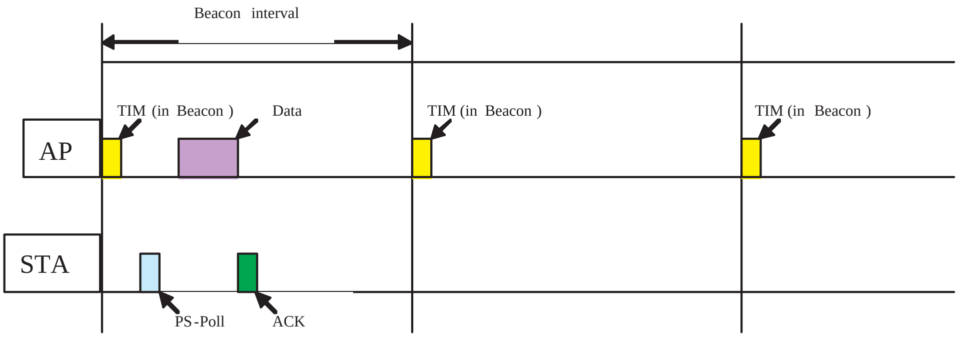

Figure 1: IEEE 802.11 power saving mode.

number of stations waking up at beacon interval t. Then, $n(t)$ can be found by

$$n(t) = \sum_{i \in S} w_i(t)$$

where set S includes all sleeping stations. Table 1 shows a sequence of $w_i(t)$ for each station and total number of wakeup stations $n(t)$, $t = 1, 2 \ldots, 18$.

t	1	2	3	4	5	6	7	8	9	10	11	12	13	14	15	16	17	18
$w_A(t)$	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1
$w_B(t)$	0	1	0	1	0	1	0	1	0	1	0	1	0	1	0	1	0	1
$w_C(t)$	1	0	0	1	0	0	1	0	0	1	0	0	1	0	0	1	0	0
$w_D(t)$	1	0	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0
$w_E(t)$	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0	1	0
$w_F(t)$	0	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0	1
$n(t)$	3	2	1	3	2	3	3	2	1	3	2	3	3	2	1	3	2	3

Table 1: A sequence of $w_i(t)$ and $n(t)$

Assume that a station J enters the power saving mode at beacon interval 3 and its listen interval is 3. Table 2 shows a sequence of $w_i(t)$ and $n(t)$, $t = 1, 2, \ldots, 18$ after station J joins in power saving mode. Note that there are 4 stations that will wake up at beacon intervals 6, 12 and 18. If these stations (i.e., stations, A, B, F and J) find that the AP has buffered data for them at beacon intervals 6, 12 or 18, they may suffer the collisions for sending PS-Poll to retrieve the buffered data. Note that collision causes retransmission and thus the station consumes more power for retrieving the buffered data. If we can arrange the first wakeup time for station J, the maximum of $n(t)$ can be reduced. This can save mobile stations' power. For example, if we can schedule the first wakeup time for station J at beacon interval 5, the maximum number of contending stations will be reduced to 3 (see Table 3).

t	1	2	3	4	5	6	7	8	9	10	11	12	13	14	15	16	17	18
$w_A(t)$	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1
$w_B(t)$	0	1	0	1	0	1	0	1	0	1	0	1	0	1	0	1	0	1
$w_C(t)$	1	0	0	1	0	0	1	0	0	1	0	0	1	0	0	1	0	0
$w_D(t)$	1	0	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0
$w_E(t)$	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0	1	0
$w_F(t)$	0	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0	1
$w_J(t)$	-	-	-	0	0	1	0	0	1	0	0	1	0	0	1	0	0	1
$n(t)$	3	2	1	3	2	4	3	2	2	3	2	4	3	2	2	3	2	4

Table 2: A sequence of $w_i(t)$ and $n(t)$ after station J joins in power saving mode.

t	1	2	3	4	5	6	7	8	9	10	11	12	13	14	15	16	17	18
$w_A(t)$	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1	1
$w_B(t)$	0	1	0	1	0	1	0	1	0	1	0	1	0	1	0	1	0	1
$w_C(t)$	1	0	0	1	0	0	1	0	0	1	0	0	1	0	0	1	0	0
$w_D(t)$	1	0	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0
$w_E(t)$	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0	1	0
$w_F(t)$	0	0	0	0	0	1	0	0	0	0	0	1	0	0	0	0	0	1
$w_J(t)$	-	-	-	0	1	0	0	1	0	0	1	0	0	1	0	0	1	0
$n(t)$	3	2	1	3	3	3	3	3	1	3	3	3	3	3	1	3	3	3

Table 3: A sequence of $w_i(t)$ and $n(t)$ for station J with first wakeup time $t = 5$.

In order to trace the wakeup time of each station, AP needs to maintain a wakeup counter, denoted as $c_i(t)$, for each sleeping station i. The $c_i(t)$ indicates remaining beacon intervals that station i will wake up. Initially, AP sets $c_i(t) = \ell_i - 1$ for station i wherever station i enters the sleeping mode. Where ℓ_i is the listen interval of station i. The counter $c_i(t)$ will be decreased by one after beacon frame is transmitted. A station i wakes up if its $c_i(t)$ becomes zero. After that, the counter will reset to $c_i(t) = \ell_i - 1$ for further counting down. Thus, after transmitting a beacon frame, AP sets counter $c_i(t+1), i \in S$, for beacon interval $t+1$ as follows:

$$c_i(t+1) = \begin{cases} c_i(t) - 1, & \text{if } c_i(t) \neq 0 \\ \ell_i - 1, & \text{if } c_i(t) = 0 \end{cases}$$

In general, if a mobile station i has wakeup counter $c_i(t)$ and listen interval ℓ_i at beacon interval t, AP can find $w_i(t+k)$ for beacon interval $t+k$ by

$$w_i(t+k) = \begin{cases} 1 & \text{if } k \bmod \ell_i = c_i(t) \\ 0 & \text{otherwise} \end{cases}$$

where $k = 1, 2, \ldots$. Then, $n(t+k)$ can be found by computing $n(t+k) = \sum_{i \in S} w_i(t+k)$ for $k = 1, 2, \ldots$ where S includes all sleeping stations.

The wakeup scheduling problem (WSP) considered in this paper can be stated formally as follows: Given a set of sleeping station S at beacon interval t consisting m stations and each of the station i having wakeup counter $c_i(t)$ and ℓ_i, for a new sleeping station j assign an initial value to its $c_j(t)$ such that the maximum value of $n(t+k) = \sum_{i \in S \cup \{j\}} w_i(t+k), k = 1, 2, \ldots$ is minimized.

By observing the sequence of $n(t)$ in Table 1, we find that a pattern repeats every six beacon intervals, e.g., $(n(1), n(2), \ldots, n(6)) = (n(7), n(8), \ldots, n(12)) = (n(13), n(14), \ldots, n(18)) = (3, 2, 1, 3, 2, 3)$. The length of this repeating pattern, r, can be found by computing the least common multiple (lcm) of listen-interval $\ell_i, i \in S$. For example, the listen intervals of stations, A, B, C, D, E, and F are 1, 2, 3, 6, 6, and 6, respectively, in Table 1. Then

$$r = lcm\{1, 2, 3, 6, 6, 6\} = 6$$

Thus, $n^* = \max\{n(t+1), n(t+2), \ldots, n(t+r)\} = \max\{n(t+k) | k = 1, 2, \ldots\}$. Now, we want to add a new sleeping station j with listen interval ℓ_j to the sleeping set S with repeating pattern size r and assign an initial value to $c_j(t)$. A stepwise solving method for WSP problem is given as follows.

1. Find $r = lcm\{\ell_j, r\}$ and a sequence of total number of wakeup stations $(n(t+1), n(t+2), \ldots, n(t+r))$ for the first r intervals for sleeping station set $S \cup \{j\}$.

2. For $i = \ell_j - 1, \ldots, 1, 0$, perform the following operations:

   (a) Set $c_j(t) = i$ and find $(w_j(t+1), w_j(t+2), \ldots, w_j(t+r))$;

   (b) Set $(n(t+1), n(t+2), \ldots, n(t+r)) = (n(t+1), n(t+2), \ldots, n(t+r)) + (w_j(t+1), w_j(t+2), \ldots, w_j(t+r))$;

   (c) Find $n_i = \max\{n(t+1), n(t+2), \ldots, n(t+r)\}$.

3. Find $n^* = \min\{n_i | i = \ell_j - 1, \ell_j - 2, \ldots, 0\}$, say $n^* = n_k$, and thus set $c_j(t) = k$.

For example, six stations with $r = 6$ as given in Table 1, station J with $\ell_J = 3$ enters the sleeping mode. The AP applies the above solving method to determine the initial value of counter $c_J(t)$ for station J as follows.

1. $r = lcm\{3, 6\} = 6$ and $(n(t+1), n(t+2), \ldots, n(t+6)) = (3, 2, 1, 3, 2, 3)$

2. $i = 2$:

   (a) Set $c_j(t) = 2$ and find $(w_j(t+1), w_j(t+2), \ldots, w_j(t+6)) = (0, 0, 1, 0, 0, 1)$;

   (b) Set $(n(t+1), n(t+2), \ldots, n(t+6)) = (3, 2, 1, 3, 2, 3) + (0, 0, 1, 0, 0, 1) = (3, 2, 2, 3, 2, 4)$;

   (c) Find $n_2 = \max\{3, 2, 2, 3, 2, 4\} = 4$.

   $i = 1$:

   (a) Set $c_j(t) = 1$ and find $(w_j(t+1), w_j(t+2), \ldots, w_j(t+6)) = (0, 1, 0, 0, 1, 0)$;

   (b) Set $(n(t+1), n(t+2), \ldots, n(t+6)) = (3, 2, 1, 3, 2, 3) + (0, 1, 0, 0, 1, 0) = (3, 3, 1, 3, 3, 3)$;

   (c) Find $n_1 = \max\{3, 3, 1, 3, 3, 3\} = 3$.

$i = 0$:

(a) Set $c_j(t) = 0$ and find $(w_j(t{+}1), w_j(t{+}2), \ldots, w_j(t{+}6)) = (1, 0, 0, 1, 0, 0)$;

(b) Set $(n(t+1), n(t+2), \ldots, n(t+6)) = (3, 2, 1, 3, 2, 3) + (1, 0, 0, 1, 0, 0) = (4, 2, 1, 4, 2, 3)$;

(c) Find $n_0 = \max\{4, 2, 1, 4, 2, 3\} = 4$.

3. Find $n^* = \min\{4, 3, 4\} = 3$, i.e., $n^* = n_1$, and thus set $c_j(t) = 1$.

Note that according to IEEE 802.11 standard, if mobile station j has no data to send, it can send a Null data frame with Power Management bit set to AP. The AP begins buffering frames and sends an ACK frame to the station after receiving the Null data frame. We can just modify this step to incorporate our wakeup scheduling in IEEE 802.11 standard as follows: The AP begins buffering frames, determines $c_j(t)$ and sends an ACK frame with $c_j(t)$ value to station j after receiving the Null data frame. Then, the station j sets its wakeup counter to $c_j(t)$ and enters the sleeping mode.

4 Contention Avoidance Traffic Scheduling

In the previous section, we arrange stations' wakeup times so that the number of wakeup stations is balanced. In this section, we consider how to inform stations that frames are buffered such that the contention is avoided. Three different access scheduling mechanisms are proposed for the contention avoidance problem. In the first mechanism, only one wakeup station is scheduled to access the buffered data in a beacon interval by marking one bit in TIM. The second and third mechanisms schedule multiple wakeup stations to get back their buffered data within a beacon interval. The access sequence within the beacon interval is according to their AIDs and the length of queuing data.

4.1 Multiple Wakeups Single Access

One of simple ways to avoid contention is that we only choose a station to inform it that AP has its buffered frames at each beacon interval. So there is no contention problem of sending PS-Poll frame to get back its buffered data. Let $S_w(t)$ be the set including all stations waking up at beacon interval t. That is,

$$S_w(t) = \{i | i \in S, c_i(t) = 0\}$$

where S is the set including all sleeping stations. Let $S_b(t)$ be the set including all stations that frames are buffered in AP at beacon interval t. Thus, we can choose a station, say station v, from set $S_w(t) \cap S_b(t)$ with a largest listen interval ℓ_v to inform that the AP has buffered frames for it.

It is possible that there may exist some stations with small listen interval in set $S_w(t) \cap S_b(t)$ and they are never chosen by AP. To avoid such a case, we associate each station v in $S_w(t) \cap S_b(t)$ with an age, denoted as a_v. Initially, the age of each station is set to zero. For each beacon interval, if a station in set $S_w(t) \cap S_b(t)$ is not selected to inform, AP increases its age by one; otherwise, AP sets its age to zero. Thus, AP can choose a station, say station v, from set $S_w(t) \cap S_b(t)$ with a largest value of $\ell_v + a_v$ to inform.

Figure 2 shows an example of this mechanism. Consider that there are four stations, A, B, C, and D, with listen interval $(\ell_A, \ell_B, \ell_C, \ell_D) = (2, 2, 3, 1)$. Packet arrival rate of each station is one frame per beacon interval. In beacon interval t, stations A, C, and D wake up in which station C, has maximum $p_C = \ell_C + a_C = 3 + 0$, is indicated in TIM to inform it that the AP has buffered its data. Stations A and D are deferred to their next wakeup beacon intervals. The AP sets ages $a_A = a_A + 1$ and $a_D = a_D + 1$. In the beacon interval $t + 1$, stations B and D wake up. Because $\ell_B + a_B = \ell_C + a_C = 2$, the AP selects station B, arbitrarily, to inform it has buffered frames. Similarly, station A is chosen to inform in beacon interval $t + 2$. At beacon interval $t + 3$, $\ell_B + a_B < \ell_C + a_C < \ell_D + a_D$ and thus station D is chosen to inform.

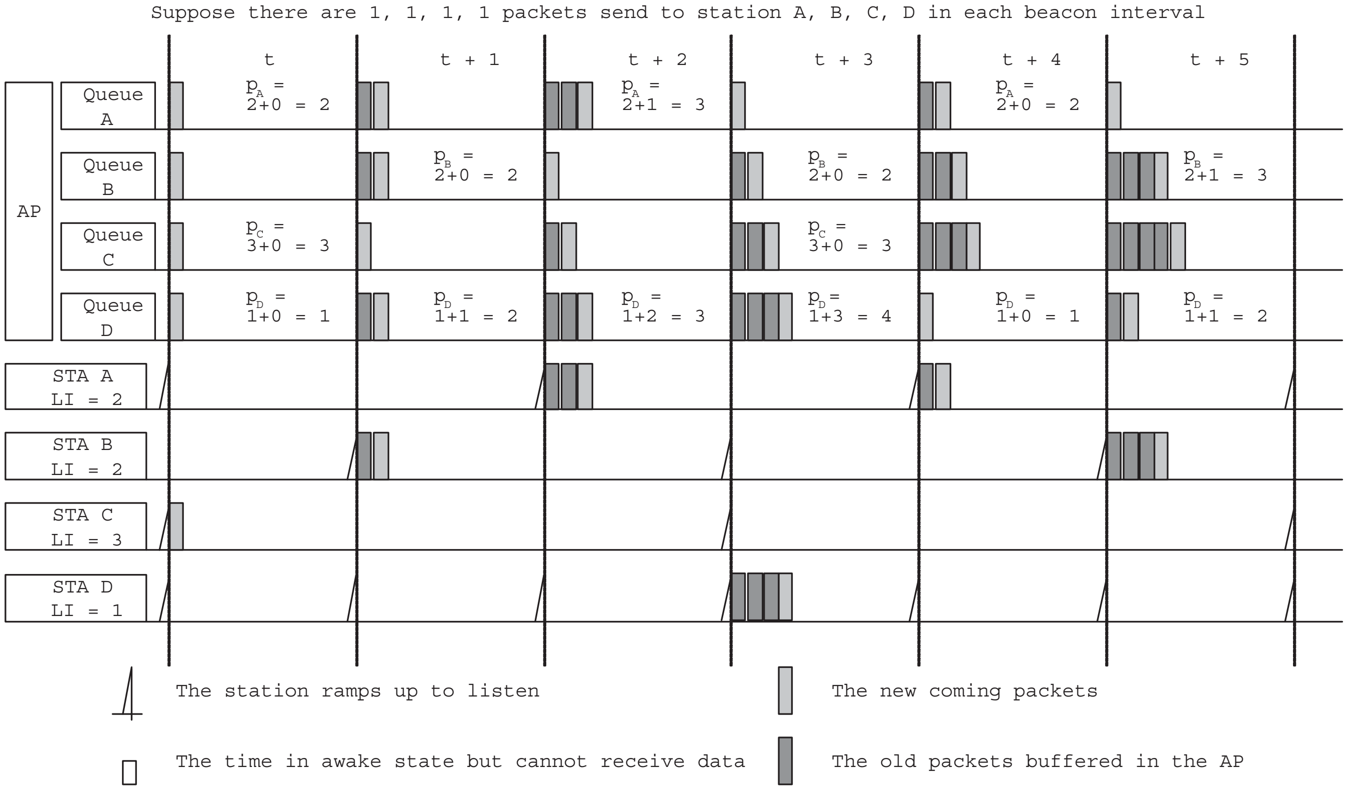

Figure 2: Multiple Wakeup Single Access

4.2 Multiple Wakeups Multiple Accesses

Although the multiple wakeups single access mechanism avoids the contention among stations, it may lower the bandwidth utilization and increase the transmission delay. However, the AP knows how many frames it has buffered in queue, transmission rate and the length of beacon interval. Thus, the AP can determine how many frames it can transmit in a beacon interval and schedule the buffered frame by means of announcing the TIM. In the following, we give two methods to arrange the access sequences of stations.

4.2.1 The smallest AID first

In order to control the traffic load in a beacon interval, AP selects a set of stations with an appropriate size from $S_w(t) \cap S_b(t)$ to inform them to retrieve the data.

That is the total amount of buffered frames of selected stations should be less than the capacity of a listen interval. This can avoid a station that awakes within whole beacon interval but can not get back its buffered data. Next, we modify the power management scheme of 802.11 WLAN such that a station to retrieve the buffered frame according to the sequence of AID marked in TIM. That is, the station with smallest AID among the selected stations sends PS-Poll frame to retrieve buffered data first.

Figure 3 shows an example of the AID sequence method. There are four stations, A, B, C, and D with listen interval $(\ell_A, \ell_B, \ell_C, \ell_D) = (2, 1, 3, 2)$ with under the service of an AP. Their corresponding AIDs are 1, 2, 3, and 4 for stations A, B, C, and D, respectively. Suppose packet arrival rates of stations A, B, C, and D are 2, 2, 1, and 2 per beacon interval. The maximum size that AP can transfer to stations in a beacon interval is 8 frames. In beacon interval t, all of four stations wake up. Because the number of buffered frames is $2 + 2 + 1 + 2 = 7$ $(7 < 8)$, the AP marks AIDs 1, 2, 3, and 4 in the TIM. The stations check the TIM in beacon frame. They learn that 4 stations will send PS-Poll to retrieve their buffered frames and every station knows which station precedes it in access sequence. For example, station C has to wait stations A and B finishing their access. In beacon interval $t+2$, $S_w(t+2) \cap S_b(t+2) = \{A, B, D\}$ and the number of frames buffered for stations, A, B and D is 10 $(10 > 8)$. Thus, based on the values of p_A and p_D, the AP selects stations A and D to inform them to retrieve the buffered data.

4.2.2 The smallest queue length first

Instead of the smallest AID first, the AP can arrange the access sequence for the selected stations according to their associated queue lengths. The station with smallest queue length receives a highest precedence and thus it can have a longer sleeping time. In this method, we need to add an information element, describes the access sequence, as a component of the beacon frame. The station checks this

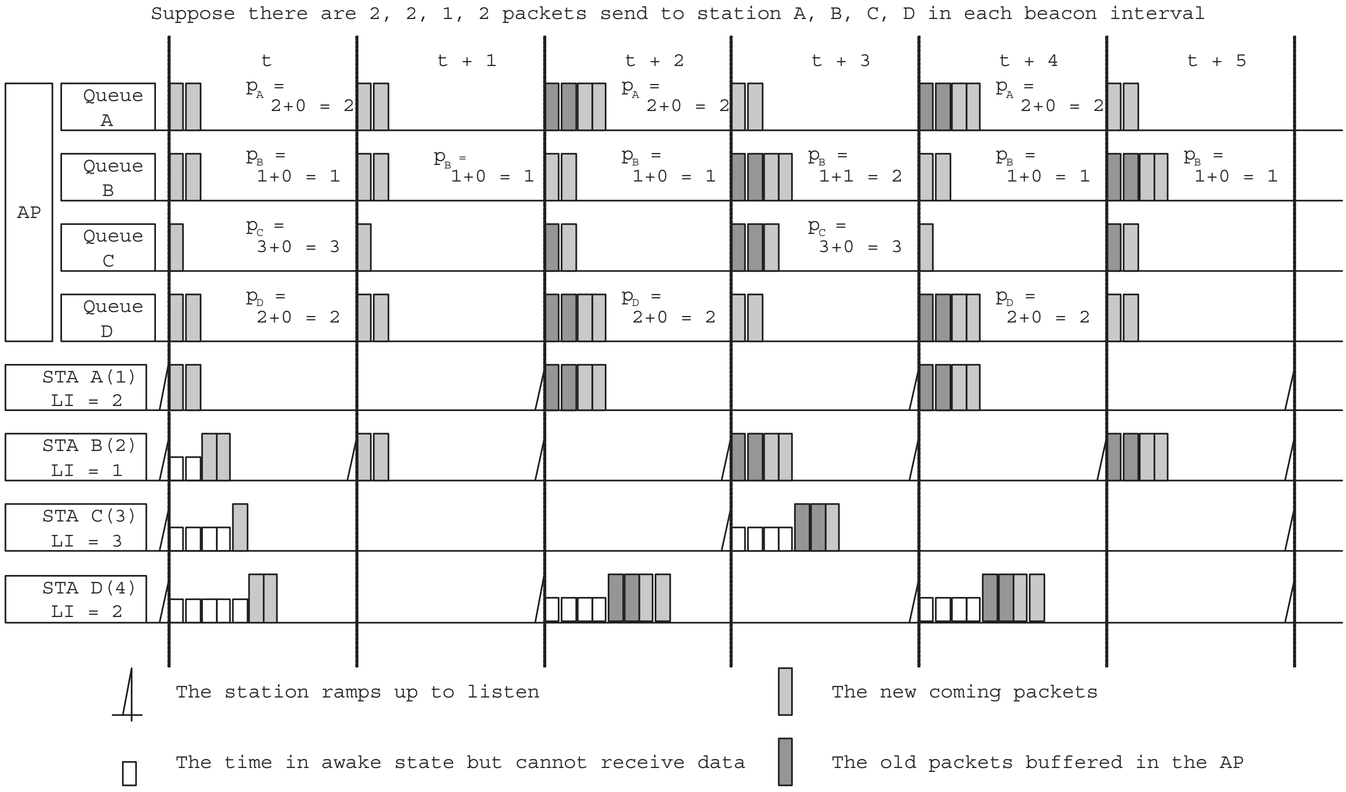

Figure 3: An example of the smallest AID first method

information element for the access sequence.

Figure 4 shows an example of the smallest queue length first method. There are three stations, A, B, and C with listen interval $(\ell_A, \ell_B, \ell_C) = (2, 1, 2)$ with under the service of an AP. Suppose packet arrival rates of stations A, B, and C are 2, 2, and 1 frames per beacon interval. The maximum size that AP can transfer to stations in a beacon interval is assumed to be 8 frames. In beacon interval t, all of the three stations wake up. Because the number of buffered frames are $2 + 2 + 1 = 5$ $(5 < 8)$, the AP marks AIDs 1, 2, and 3 in the TIM and adds the access sequence C, A, and B in the beacon frame. In beacon interval $t + 2$, the access sequence is stations C, B, and A. Note that if two stations have same queue length, AP uses their p_v values to break the tie.

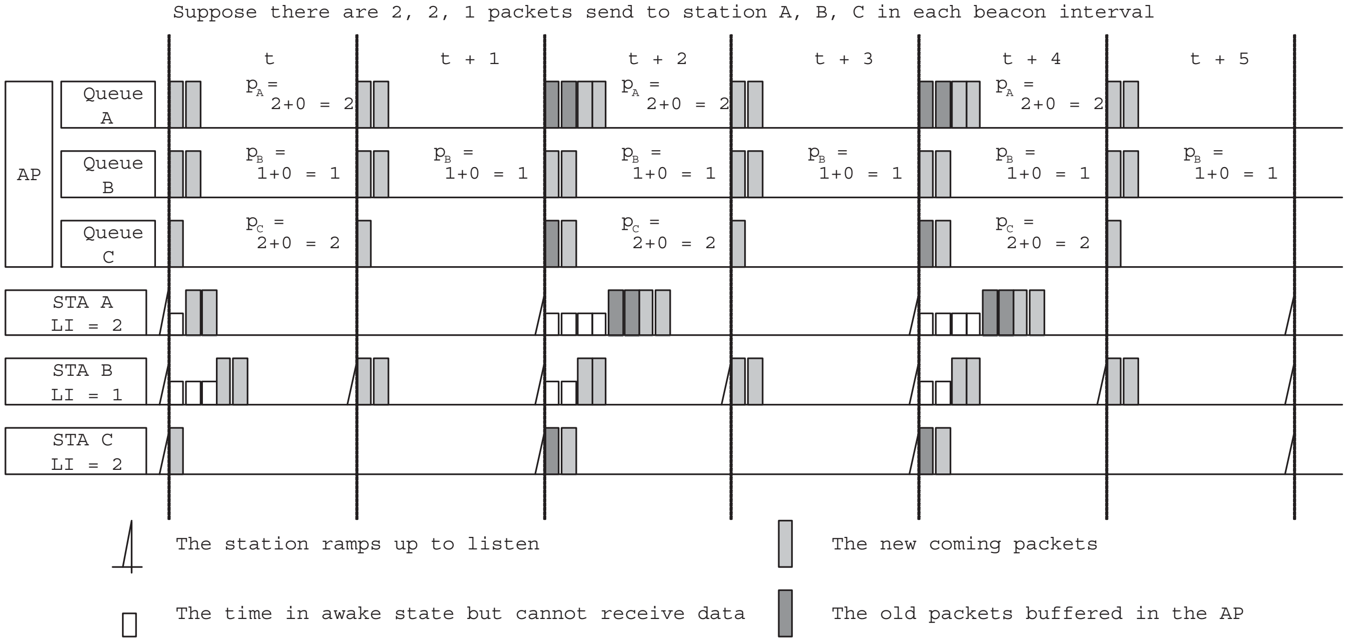

Figure 4: An example of the smallest queue length first method

5 Simulation and results

5.1 Performance metrics and environment setup

In this section, we show the performance analysis for the proposed schemes:

1. Load-Aware Wakeup Scheduling (LAWS);
2. LAWS with Multiple Wakeups Single Access (LAWS+MWSA);
3. LAWS with multiple wakeups multiple access and the Smallest AID First (LAWS+SAF);
4. LAWS with multiple wakeups multiple access and the Smallest Queue Length First (LAWS+SQLF).

Note that all four schemes are enhancements of the 802.11 PS mode. The LAWS arranges station's wakeup time. The MWSA, SAF and SQLF schemes can be used by AP to schedule the access sequence by marking the bits in TIM. We compare their performances again pure 802.11 PS mode by simulation. The performance metrics are given as follows:

1. Average sleeping time of the station: This measure is the duration that a station stays in the sleeping mode. If a scheme can make stations stay more time in sleeping, then stations will save more power.

2. Average throughput: This value shows the total amount of data successfully transmitting per second. If AP can efficiently schedule and distribute the access of its serving stations, it will have higher data throughput.

3. Average latency of a successful transmission: The latency is defined as the time duration starting while a packet is issued and buffered at AP and ending when the target station returns the acknowledge. An AP with a good scheduling scheme will make the latency as small as possible. Thus, the resources required for buffering data can be reduced.

Our simulation uses an IEEE 802.11b wireless communication module with 11Mbps data rate. An AP can serve at most 30 stations. Each station will randomly set 1 to 5 beacon intervals as its listen interval size and its packet arrival rate is 3 packets per beacon interval. Packet size in our simulation is fixed and set to 1KB. Communication channel assumes to be clear and symmetric. The total simulation time is 3 minutes. The details of other simulation configurations such as header length, and inter-frame spaces (IFS) are listed in table 4. Simulation results will compare the IEEE 802.11 PS mode with the proposed LAWS, LAWS+MWSA, LAWS+SAF, and LAWS+SQLF schemes.

5.2 Results and Discussion

Figure 5 shows the relation between average sleeping time and number of stations. Considering contention-based schemes, LAWS can have more sleeping time than IEEE 802.11 PS mode in any size of stations. By using LAWS+MWSA, LAWS+SAF, and LAWS+SQLF schemes to reduce the contention within a beacon interval, stations can have more time on staying in sleeping than LAWS and IEEE

Data Rates	11Mbps
MAC header	28 bytes
IP header	20 bytes
UDP header	20 bytes
Beacon frame	28 bytes
ACK frame	14 bytes
PS-Poll frame	14 bytes
SIFS	0.00001 second
DIFS	0.00005 second
Slot time	0.00002 second
Beacon interval	0.1 second

Table 4: Detail simulation configurations.

802.11 PS mode. In this figure, it seems that LAWS+MWSA has better sleeping time than both LAWS+SAF and LAWS+SQLF. However, we will find in Figure 7 that it trades the transmission latency with the sleeping time.

Figure 6 shows the average throughput for each scheme. From this figure, we can explicitly find that the throughput of IEEE 802.11 PS mode falls down when station number is greater than 20. However, our proposed schemes, LAWS, LAWS+MWSA, LAWS+SAF, and LAWS+SQLF, are not influenced as number of station increases. This is because our schemes can efficiently avoid the data collision among the stations.

In Figure 7, we show the average latency of a successful transmission for each scheme. For LAWS, LAWS+SAF, and LAWS+SQLF, all of their latency is smaller than 0.3 second and increase slowly as number of stations grows. Because only one station is indicated within a beacon interval, the latency of LAWS+MWSA scheme is longer than the other proposed schemes. The pure IEEE 802.11 PS mode, however, will suffer the worst latency while number of stations increases.

Finally, Figure 8 shows the improving rate of sleeping time for each proposed scheme (compared to pure IEEE 802.11 PS mode). The improving rate R_i of scheme i is defined as $R_i = \frac{S_i - S_0}{S_0} \times 100\%$, where S_0 and S_i are the average sleep-

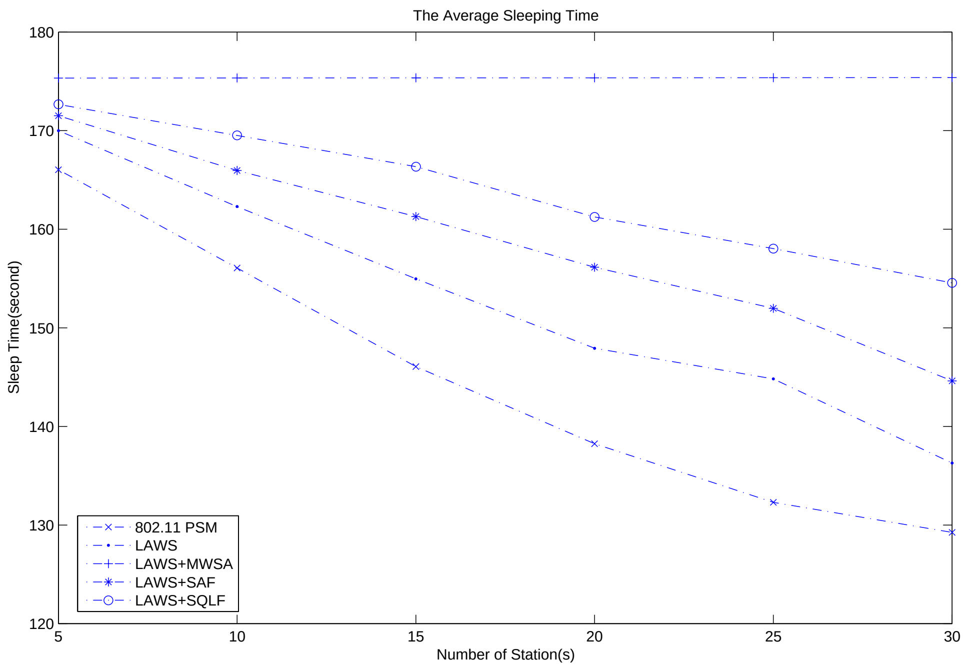

Figure 5: The average sleeping time.

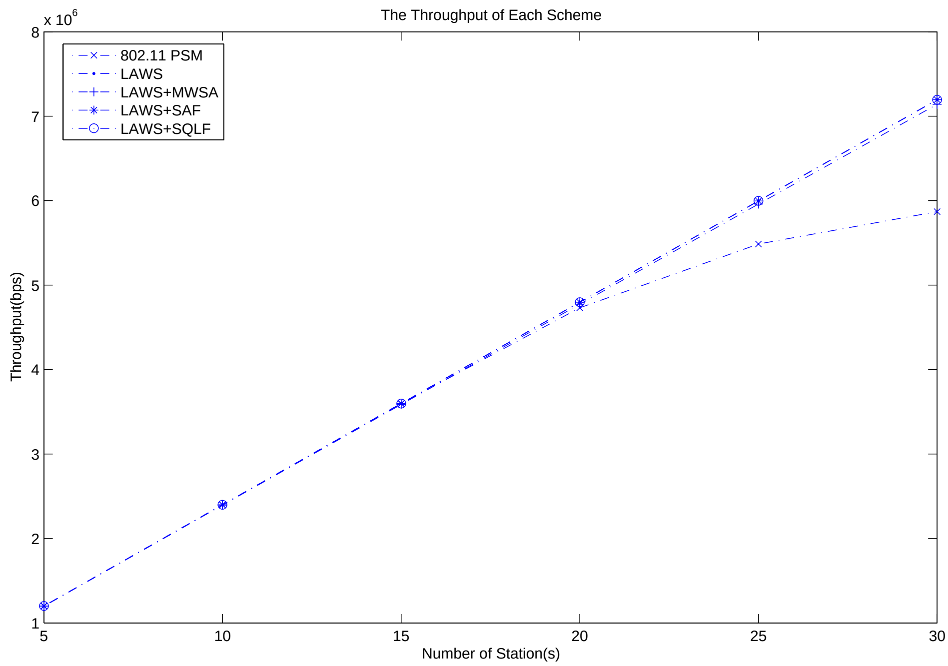

Figure 6: The average throughput for each scheme.

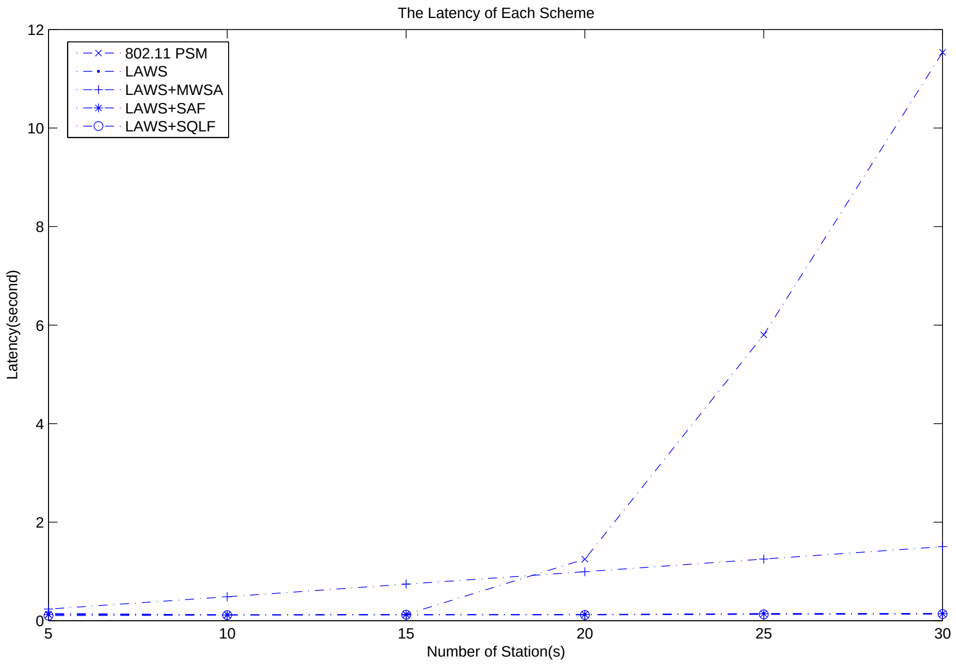

Figure 7: The latency of a successful transmission for each scheme.

ing times for pure 802.11 PS mode and the proposed scheme i, respectively. By efficiently scheduling the wakeup time of sleeping stations, the sleeping duration of LAWS, LAWS+MWSA, LAWS+SAF, and LAWS+SQLF schemes can be improved significantly.

6 Conclusion

In this paper, we propose a load-aware wakeup schedule scheme for infrastructure mode of IEEE 802.11 WLANs. The LAWS scheme balances the number of wakeup stations in each beacon interval to reduce the amount of contention stations. For avoiding the contention, MWSA, SAF, and SQLF scheme are used to arrange the access sequence of the wakeup stations within a beacon interval. Simulation results show that comparing to 802.11 PS-mode, the proposed LAWS, MWSA, SAF, and SQLF schemes can efficiently improve the sleeping duration of each station, average throughput and transmission delay.

The following two issues should be considered in the implementation of the

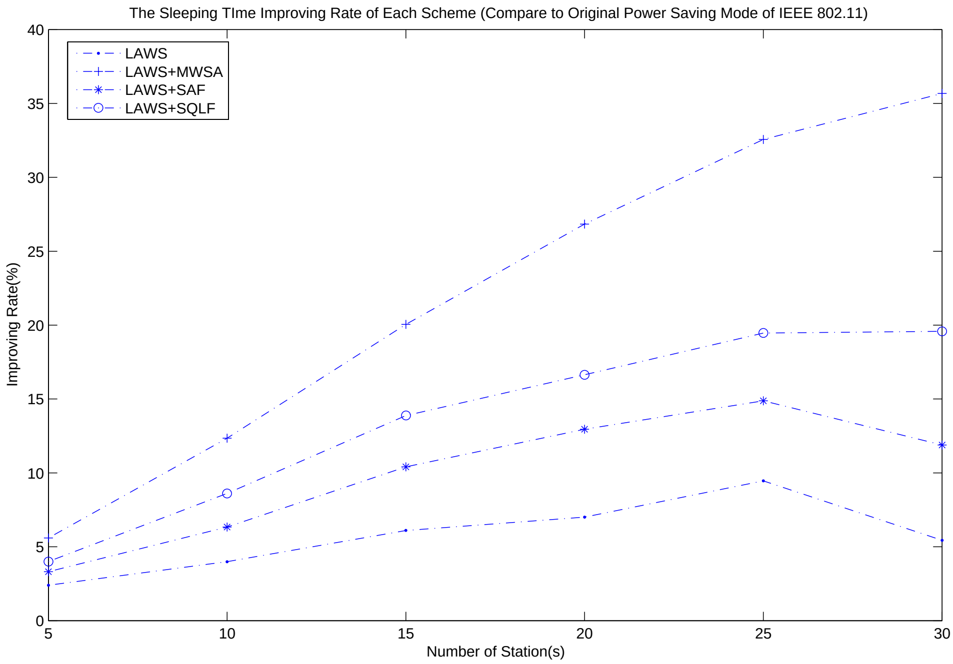

Figure 8: The improving rate of sleeping time (compared to PS mode of IEEE 802.11).

proposed schemes:

1. An *aging function* should be implemented in the AP to determine when buffered frames are old enough to be discarded.

2. If the mobile station misses the beacon, it should remain awake until it receives the next beacon. The mobile station checks the beacon frame. If the bit corresponding to its AID is set to zero in the TIM, or else it has retrieved all buffered frames, the mobile station can resume the sleeping mode by asking AP for a new wakeup counter $c_j(t)$. In the LAWS+SAF and LAWS+SQLF schemes, the mobile station misses the beacon can not show up to retrieve the buffered data in its turn. The next station in the access sequence can send PS-Poll frames to get back its buffered data if it finds that the medium has been idle for longer than the Distributed coordination function Inter-Frame Space (DIFS).

Finally, finding a shorter repeating pattern for LAWS scheme or extending

LAWS scheme to wireless ad hoc networks might be interesting for possible future work.

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A Weighted Multilateration Positioning Method for Wireless Sensor Networks*

Yu-He Gau, Hung-Chi Chu, and Rong-Hong Jan†
Department of Computer and Information Science
National Chiao Tung University, Hsinchu, 30050, Taiwan

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***Abstract*— Localization problem is one of the most important research issues for wireless sensor networks (WSNs). In this paper, we present a two-phase localization algorithm for WSNs. In the first phase, each sensor node obtains its initial position by DV-hop method. In the second phase, each sensor node gathers the locations and distances to its neighbors, updates and exchanges these location information periodically and then operates the multilateration with different weight values. The simulation result shows that the average position error of the proposed two-phase method is less than 20% of the radio range, and the number of sensor nodes which can be located is larger than 70% of total nodes for a network with low density.**

***Index Terms*— Multilateration positioning method, Wireless sensor networks**

I. Introduction

The topology of wireless sensor network (WSN) is similar to ad-hoc network in which each sensor node in the network should communicate and cooperate with its neighbors to achieve the goal of task. Each node collects, stores, and processes the sensed data, such as temperature, brightness, sound, and so on, and then communicates with neighboring nodes to provide the environmental observation.

In order to make the sensed data more useful, the location of sensor nodes should be determined. There are several location determination methods have been presented [1-15]. In general, the location methods can be divided into two main classes, one is centralized system and the other is distributed system. The centralized system has a central location server receiving the location query, calculating the location information, and replying it back to the query node. In contrast, the distributed system, each node utilizes the location algorithm to calculate its position by itself. The distributed system is more feasible than the centralized system because of the following reasons: First, the location server is a bottleneck of the centralized system and nodes near the location server consume more energy in forwarding location information. Second, system stability depends on the communication links of central location server. If these links are failure, sensor nodes fail to determine their locations.

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[2]Corresponding Author. Fax: 886-3-5721490; e-mail: rhjan@cis.nctu.edu.tw.

One of the most popular distributed location methods is Global Positioning System (GPS)[1] that has been shown to be an integrated part of modern navigation. However, it is not suitable for all sensor nodes because of the limitation of sensor node in size, cost, electric power and computational power. Besides, GPS signal is degraded by the environment or jamming. Especially in indoor environments, GPS signal would be blocked and almost unavailable.

In this paper, we develop a distributed location method for sensor nodes with some *beacon nodes*. The *beacon node*, (or called as *beacon* in short) is a sensor node that knows its location. This paper presents a two-phase location method for WSN with a set of beacons. In the first phase, each node estimates its initial position by a rough DV-hop method[2]. In the second phase, each sensor node gathers the location information that includes the locations and distances of its neighbors. And then it updates and exchanges the location information periodically, and operates the multilateration with different weight values. The weight is defined as an error function of positions and distances to show the effect of data accuracy. The node having a higher weight value implies that the position data of node is more accurate. However, the multilateration technique does not work well if the number of available neighbors' location is less then three. Thus, in this paper, we also propose a method to deal with the multilateration with two neighbors. The simulation results show that the average position error is noticeable.

The remainder of this paper is organized as follows. In section 2, the related work of existing location method will be introduced. The weighted multilateration algorithm and the simulation results are shown in sections 3 and 4, respectively. Finally, the conclusion and future work are given in section 5.

II. Related work

For WSN applications, localization is one of the important issues. In [3], a survey paper notes that early classical location systems such as RADAR [4], Cricket [5], and Active Badge [6] may not suitable for sensor networks. The limitation of sensor nodes in size, cost, and power consumption should be considered for localization of WSN.

Based on the operation model, the positioning system can be divided into two main classes: anchor-based and anchor-free.

Anchor node, a special sensor node, can obtain its location, floods its location information to the network providing location information for other nodes and has unlimited power.

In anchor-based location system, it utilizes some anchor nodes to achieve location determination. In [7] and [8], they use Multidimensional Scaling (MDS) to get global relative coordinate system, and then map this system into absolute coordinate system via anchor node by self-designed location algorithm. One of the popular anchor-based location systems is GPS[1]. It relies on the 24 satellites that orbit around the earth to transmit precise velocity, latitude, longitude, and altitude information to the GPS receiver. Those 24 satellites are the anchor nodes that we mentioned in this paper. Ssu et al.[9] proposed a localization method with mobile anchor nodes. Each sensor node records the track of mobile anchor with lifetime and analyzes them to find out node's location. In this method, mobile anchor nodes can reduce the number of anchor nodes but it should guarantee that all sensor nodes can obtain enough information that transmitted from mobile anchor node to perform this localization method. In [10], they utilize the radio transmission signal with overlapping to define several unique region and process the location information that is transmitted by anchor nodes to obtain nodes' location. In [11], the localization method is divided into two phase. In the first phase, it utilizes the Hop-TERRAIN algorithm that is similar to DV-hop[2] to determine a coarse location of sensor node. The Hop-TERRAIN algorithm is worked by location information that anchor node or sensor node broadcast. In the second phase, it refines the location that was obtained in the first phase.

In contrast, the location system without anchor nodes is called as anchor-free location system. In [12] and [13], nodes exchange local distance information with others to generate their relative coordinates. The transformation from relative coordinates to absolute coordinates was handed over to certain post-process methods. In [14], sensor nodes are put into n group and broadcast its ID and group ID. Using the information of node ID and group ID, it can localize sensor nodes. In [15], each node discovers its neighbors that are within its transmission range and estimates their ranges. It consists three phases: node discovery, range estimation and cluster formation is performed in the first phase; in the second phase, a local map is built at each gateway node (cluster leader) using the range measurements; gateway node collaborates to obtain a global map of the network in the last phase.

III. The weighted multilateration algorithm

The proposed method, called as a weighted multilateration positioning method is a distributed location method. It includes two phases. In the first phase, the DV-hop method in APS [2] is modified. Beacon nodes broadcast their position information, and then non-beacon nodes can get their initial positions. In the second phase, non-beacon nodes only exchange information with their neighbors and then estimate their location by multilateration with weighted coefficient to refine initial node positions. This is because neighbors' position and distance to neighbors have different reliability. In the following, we will present the weighted multilateration algorithm in details.

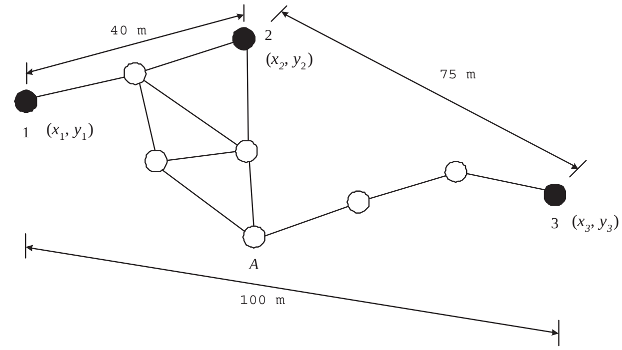

Fig. 1. The example for the first phase.

A. The first phase

The goal of the first phase is to estimate the initial positions for non-beacon nodes. This phase is modified from DV-hop method. The stepwise description is given as follows.

1) Beacon nodes broadcast their positions periodically. All non-beacon nodes will relay this beacon frame from one beacon node to the other and increase the hop count. The format of beacon frame contains the following data:
$$B = \{1, ID, (x, y), Hc, \bar{d}\},$$
where the parameter 1 indicates the system is in the first phase; ID is the identification of beacon node; (x, y) is its coordination; Hc is the hop count from beacon node to received node; and $\bar{d}$ is the average one-hop distance. Initially, $\bar{d}$ is set to zero.
2) Beacon nodes maintain Euclidean distance d_{ij} and hop count h_{kj} to other beacon nodes via receiving information in Step 1. Non-beacon nodes collect the hop count information to all beacon nodes.
3) Beacon nodes compute their average one-hop distance and then broadcast these values to non-beacon nodes.
4) Nodes receive information from at least three beacon nodes and estimate distances to beacon nodes. This is done by shortest path hop count α multiplied average one-hop distance $\bar{d}_1$. That is,
$$d_{kj} = \alpha \times \bar{d}_1,$$
where d_{kj} is the estimated distance from node k to node j.
5) Non-beacon node operates classical multilateration [16] to estimate their positions.

For example, as shown in Figure 1, nodes 1, 2, and 3 are beacon nodes and others are non-beacon nodes. The details of steps are described as follows.

Step 1: Beacon nodes 1, 2, and 3 broadcast their beacon frame which contain the following data:
$$\begin{aligned} B_1 &= \{1, 1, (x_1, y_1), 1, 0\}, \\ B_2 &= \{1, 2, (x_2, y_2), 1, 0\}, \\ B_3 &= \{1, 3, (x_3, y_3), 1, 0\}. \end{aligned}$$

Step 2: Beacon nodes 1, 2, and 3 estimate Euclidean distance to other beacon nodes, i.e. $d_{12} = d_{21} = 40m$, $d_{23} =$

$d_{32} = 75m$, and $d_{13} = d_{31} = 100m$. They also maintain the hop count between beacons, i.e. $h_{12} = h_{21} = 2$, $h_{23} = h_{32} = 5$, and $h_{13} = h_{31} = 6$. Non-beacon node A can obtain hop count information, i.e., $h_{A1} = 3$, $h_{A2} = 2$, and $h_{A3} = 3$.

Step 3: Beacon nodes compute their average one-hop distances as follows:

1) The average one-hop distance for beacon node 1 $= \frac{d_{12}+d_{13}}{h_{12}+h_{13}} = \frac{40+100}{2+6} = 17.5$,
2) The average one-hop distance for beacon node 2 $= \frac{d_{21}+d_{23}}{h_{21}+h_{23}} = \frac{40+75}{2+5} = 16.43$,
3) The average one-hop distance for beacon node 3 $= \frac{d_{31}+d_{32}}{h_{31}+h_{32}} = \frac{100+75}{6+5} = 15.91$,

Then, beacon nodes broadcast their average one-hop distances to other nodes as follows:

$$\begin{aligned} B_1 &= \{1, 1, (x_1, y_1), 1, 17.5\}, \\ B_2 &= \{1, 2, (x_2, y_2), 1, 16.43\}, \\ B_3 &= \{1, 3, (x_3, y_3), 1, 15.91\}. \end{aligned}$$

Step 4: Node A for example, it computes the average one-hop distance and then estimates distances to beacon nodes:

1) Average one-hop distance$= \frac{17.5+16.43+15.91}{3} = 16.61$;
2) Estimated distance from node A to node 1, $d_{A1} = 3 \times 16.61 = 49.83$;
3) Estimated distance from node A to node 2, $d_{A2} = 2 \times 16.61 = 33.22$;
4) Estimated distance from node A to node 3, $d_{A3} = 3 \times 16.61 = 49.83$.

Step 5: Node A operates classical multilateration to estimate its initial position:

1) Error function f_{kj} is defined as the difference between measured distance to beacon nodes, $\triangle d_{kj}$, and estimated distance, d_{kj}. For example, the error functions of node A are given as:

$$\begin{aligned} f_{A1} &= \triangle d_{A1} - d_{A1}, \\ f_{A2} &= \triangle d_{A2} - d_{A2}, \\ f_{A3} &= \triangle d_{A3} - d_{A3}, \end{aligned}$$

where $\triangle d_{A1} = \sqrt{(x_A - x_1)^2 + (y_A - y_1)^2}$, $\triangle d_{A2} = \sqrt{(x_A - x_2)^2 + (y_A - y_2)^2}$, and $\triangle d_{A3} = \sqrt{(x_A - x_3)^2 + (y_A - y_3)^2}$.

2) Apply Minimum Square Estimation (MSE) to sum of errors, i.e. $min\ f_{A1}^2 + f_{A2}^2 + f_{A3}^2$. Then, position (x_A, y_A) can be estimated.

B. The second phase

Based on the initial positions obtained in the first phase, nodes exchange location information with neighbors to refine their positions in the second phase. When nodes receive information from at least three neighbors, they can apply certain ranging technology (e.g., RSSI) to measure distances to neighbors, and then use classical multilateration to estimate their positions. However, positions and distances to neighbors have different accuracy. It is obviously that the position information from beacon nodes is more accurate than that from other estimated nodes. Similarly, according to the signal degression properties for ranging measurement, the position information from near node is more accurate than that from far node. Hence, we define a parameter, *weight*, for node i that is the production of neighbor's position weight w_{ip} and distance weight w_{id}. Position weight reflects the accuracy of neighbor's position, and distance weight reflects the accuracy of ranging measurement. The classical multilateration can be transformed into ”weighted multilateration” with given separate weight for each neighbor. The stepwise description of the second phase is given as follows.

1) A node v can receive the position data from at least three neighbors. The format of receiving beacon frame contains the following data:

$$B = \{2, ID, (x, y), weight\}.$$

It calculates the weights for its neighbors as follows.

   a) Find the position weight w_{ip} for each neighbor, where $w_{ip} = 0.1$ for non-beacon node and $w_{ip} = 1$ for beacon node.
   b) Use ranging technology (e.g. RSSI) to measure distances to each neighbor, and assign the distance weight w_d according to receiving power, where $0.1 \leq w_{id} \leq 1$. For example, if transmission power is 0.28, the receiving power and distance weight are summarized in Table I.
   c) Set the weight w_i by $w_i = w_{ip} \times w_{id}$.

2) Node v estimates its position by the weighted multilateration.
3) Node v updates its position weight by $w_{ip} = \frac{\sum w_i}{N}$ (where N is the number of neighbors), and then floods its w_{ip} and position to neighbors.
4) Repeat steps 1-3 until position error converged. In our simulation, the number of iterations needed to converge is less than ten.

TABLE I

THE RELATIONSHIP BETWEEN P_r AND w_{id}.

Receiving power P_r	w_{id}
$0.28 \times 10^{-1} \leq P_r < 0.28$	1
$0.28 \times 10^{-2} \leq P_r < 0.28 \times 10^{-1}$	0.9
$0.28 \times 10^{-3} \leq P_r < 0.28 \times 10^{-2}$	0.8
$0.28 \times 10^{-4} \leq P_r < 0.28 \times 10^{-3}$	0.7
$0.28 \times 10^{-5} \leq P_r < 0.28 \times 10^{-4}$	0.6
$0.28 \times 10^{-6} \leq P_r < 0.28 \times 10^{-5}$	0.5
$0.28 \times 10^{-7} \leq P_r < 0.28 \times 10^{-6}$	0.4
$0.28 \times 10^{-8} \leq P_r < 0.28 \times 10^{-7}$	0.3
$0.28 \times 10^{-9} \leq P_r < 0.28 \times 10^{-8}$	0.2
$P_r < 0.28 \times 10^{-9}$	0.1

As shown in Figure 2, for example, node 3 is a beacon node and others are non-beacon nodes. The stepwise description of the second phase for node 1 is summarized as follows.

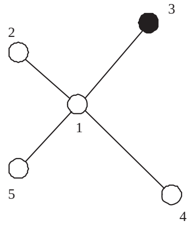

Fig. 2. The example for the second phase.

Step 1: Node 1 receives the positions data from its neighboring nodes 2, 3, 4 and 5. The format of receiving beacon frame contains the following data:

$$\begin{aligned} B_2 &= \{2, 2, (x_2, y_2), 1, 0.1\}, \\ B_3 &= \{2, 3, (x_3, y_3), 1, 1\}, \\ B_4 &= \{2, 4, (x_4, y_4), 1, 0.1\}, \\ B_5 &= \{2, 5, (x_5, y_5), 1, 0.1\}. \end{aligned}$$

(a) Find the position weight w_{ip} for each neighbor, i.e. $w_{2p} = w_{4p} = w_{5p} = 0.1$, and $w_{3p} = 1$.

(b) Use ranging technology to get the distances to its neighbors, i.e. d_{12}, d_{13}, d_{14}, and d_{15}, and obtain the distance weight, i.e. w_{2d}, w_{3d}, w_{4d}, and w_{5d}.

(c) Set the weights $w_2 = w_{2d} \cdot w_{2p}$, $w_3 = w_{3d} \cdot w_{3p}$, $w_4 = w_{4d} \cdot w_{4p}$, $w_5 = w_{5d} \cdot w_{5p}$.

Step 2: The classical multilateration is modified to ”weighted multilateration”. Apply the MSE to esitmate (x_1, y_1) by $\min\ (w_2 \cdot f_{12})^2 + (w_3 \cdot f_{13})^2 + (w_4 \cdot f_{14})^2 + (w_5 \cdot f_{15})^2$, where
$f_{12} = \sqrt{(x_1 - x_2)^2 + (y_1 - y_2)^2} - d_{12}$,
$f_{13} = \sqrt{(x_1 - x_3)^2 + (y_1 - y_3)^2} - d_{13}$,
$f_{14} = \sqrt{(x_1 - x_4)^2 + (y_1 - y_4)^2} - d_{14}$, and
$f_{15} = \sqrt{(x_1 - x_5)^2 + (y_1 - y_5)^2} - d_{15}$.

Step 3: Node 1 updates its position weight as:

$$w_{1p} = \frac{w_2 + w_3 + w_4 + w_5}{4}.$$

Step 4: Repeat the steps 1-3 for T times. The repeating times, T, is decided case by case. In simulation, T was assigned to ten.

C. Boundary process

Classical multilateration is used in the proposed method to perform the localization of nodes. We know that sensor nodes are densely deployed in the network. Most of nodes have more than three neighboring nodes in the sensor network topology to apply the classical multilateration for localization. However, some sensor nodes have just only one or two neighbors in the boundary sensing field. These nodes can not be located using classical multilateration. In order to simplify the localization problem complexity of the proposed method, we assume that all sensor nodes are fixed in the network and each one of them has more than two neighboring nodes in network topology, i.e. the connectivity of node is larger than one.

A boundary node with only two neighboring nodes, node 1 and node 2, is shown in Figure 3. The estimated location of node A in the first phase that mentioned in section 3.1, is shown as node $\bar{A}$. In geometric analysis, the location of this boundary node must be adjusted in node A' or node A''. The stepwise description of the localization of boundary node with two neighbors is given as follows.

- The boundary node receives the positioning information (x_1, y_1) and (x_2, y_2) from neighboring node 1 and node 2, respectively. Using ranging technology of RSSI, the distance d_1 and d_2 can be estimated.
- According to the positioning information and the distance to neighbors, the position of node A' and node A'' are determined by the equation of circles.

$$\begin{cases} (x - x_1)^2 + (y - y_1)^2 = d_1^2 \\ (x - x_2)^2 + (y - y_2)^2 = d_2^2 \end{cases}$$

- The boundary node calculates the distance to node A' (i.e. $\triangle d_{A'}$) and to node A'' (i.e. $\triangle d_{A''}$). The location of boundary node is in node A', if $\triangle d_{A'}$ is smaller than $\triangle d_{A''}$. On the contrary, the location of boundary node is in node A'', if $\triangle d_{A'}$ is larger than $\triangle d_{A''}$.

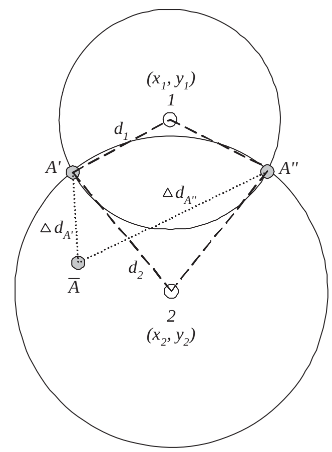

Fig. 3. A boundary node with two neighboring nodes.

IV. Simulation results

All simulations were performed in ns-2[17] environment. We discussed about connectivity and number of nodes to show the performance of the weighted multilateration algorithm. Connectivity of node represents the average number of neighboring nodes and it was affected by radio transmission range R. In order to show the effect of connectivity, we simulated various connectivity for sensor nodes about the position error and the result was shown in Figure 4. In this case, the sensing field , in which 100 sensor nodes were placed randomly, was $1000 \times 1000\ m^2$. The number of beacon nodes was 20% of total sensor nodes (i.e. beacon ratio was 20%). When the connectivity was 12, the position error was less than $50\%R$ in the first phase and was less than $10\%R$ in the second phase. The improvement of position error from the first phase to the second phase is about one-fifth when connectivity was 12.

Figure 5 shows the relationship between the number of located nodes and the connectivity of nodes. When connectivity was 4, the number of located nodes is greater than 70% both

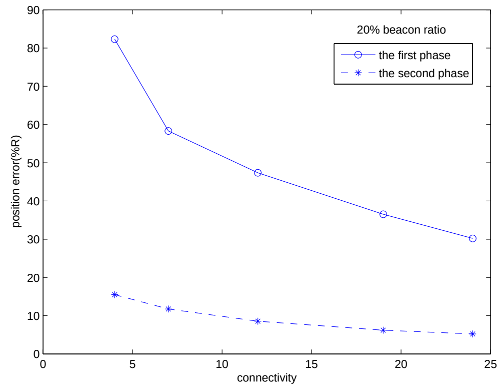

Fig. 4. The position error for various connectivity in the two phases.

in the first phase and the second phase. When connectivity is greater than 18, the number of located nodes is close to 90% in the both phases.

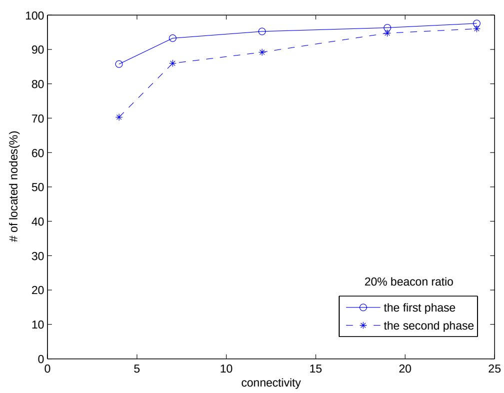

Fig. 5. The number of located nodes for various connectivity in the two phases.

The relationship about connectivity and position error with different beacon ratio for the two phases was shown in Figure 6 and 7. The position error was decrease from $200\%R$ ($40\%R$) to $50\%R$ ($10\%R$) in the first(second) phase with 5% beacon ratio. The second phase can truly refine the position error.

Compared with the robust positioning algorithm (RPA), in [11], the position error of proposed algorithm in 5%, 10%, and 20% beacon populations is better than the RPA, especially for low connectivity (see Figure 8). In contract, when the connectivity is larger than 24, the performance of the both algorithms were closed to each other. This is because that high connectivity will give more information from its neighbors to improve the both algorithms.

Next, we considered about the position error for the number of nodes in the network. The simulation results were shown in Figure 9. In this case, the sensing field is $1000 \times 1000\ m^2$. Each sensor node and beacon had the same radio range $R = 100\ m$ and they are placed randomly. In Figure 9, the position errors for both of the first and second phase are reduced as the number of nodes increases. The refinement of the second phase will keep an improvement of $50\%R$ compared to the

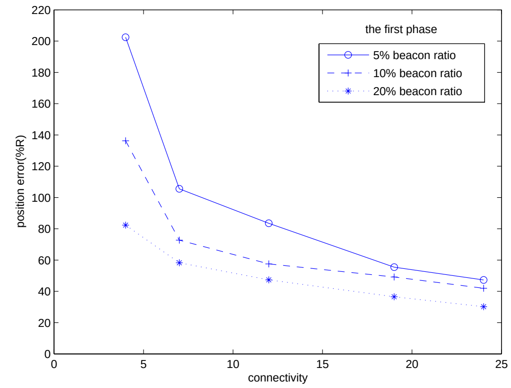

Fig. 6. The relationship between connectivity and position error for 5%, 10% and 20% beacon ratio in the first phase.

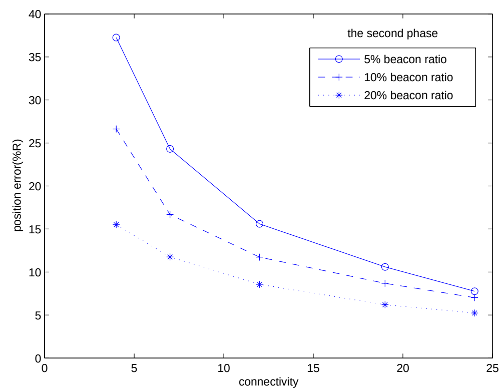

Fig. 7. The relationship between connectivity and position error for 5%, 10% and 20% beacon ratio in the second phase.

first phase.

The relationship between the number of nodes and the position error with different beacon ratio for these two phases was shown in Figures 10 and 11. The position error was decrease from $160\%R$ ($105\%R$) to $70\%R$ ($35\%R$) in the first (second) phase. The second phase can truly refine the position error from the first phase.

In real world applications, the position of beacon node can be obtained by Global positioning system (GPS) or other positioning systems that still contain some errors in position estimation. In Figure 12, we show that the effect of position error was diminished in high network density. The difference between the proposed method with and without 5% positioning error was decrease from 12.5%R to 5.1%R.

In figure 13, we present the number of iterations in weight update to converge by simulation. In the working area with 30, 70 and 100 nodes that have 20% and 30% beacon nodes for each scenario, the position error will be improved as the increase of the number of iterations. As the number of iterations is larger than 3, the position error almost keep the same value. The result shows that the position error tends to a lower bound and the overhead incurred by obtaining the weight will be restricted to T weight update for each node.

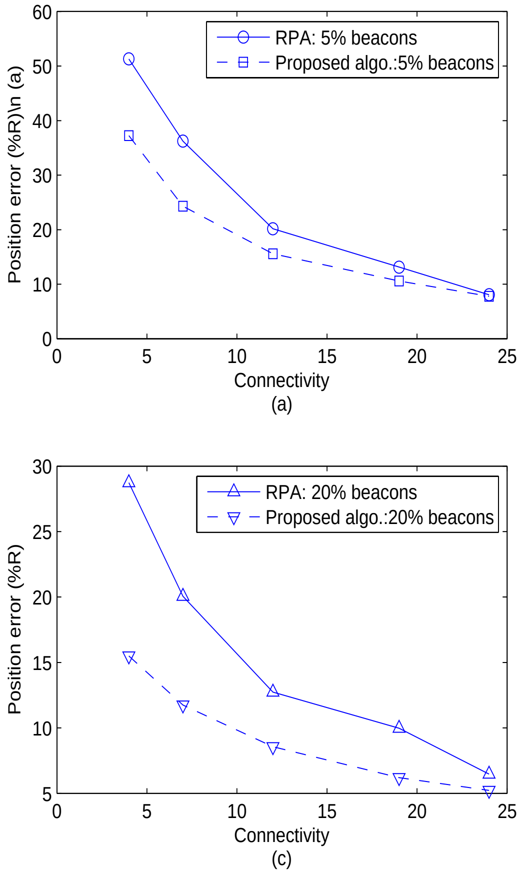

Fig. 8. Comparison of referenced algorithm (RPA) and proposed algorithm.

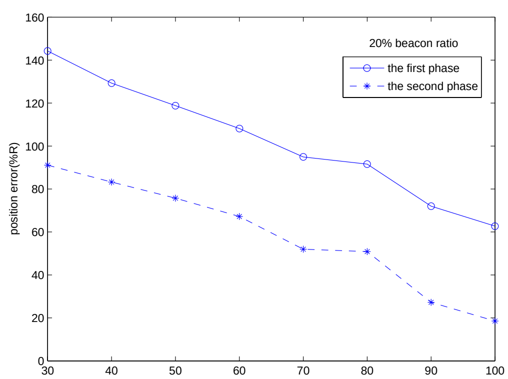

Fig. 9. The position error in the two phases with various number of nodes.

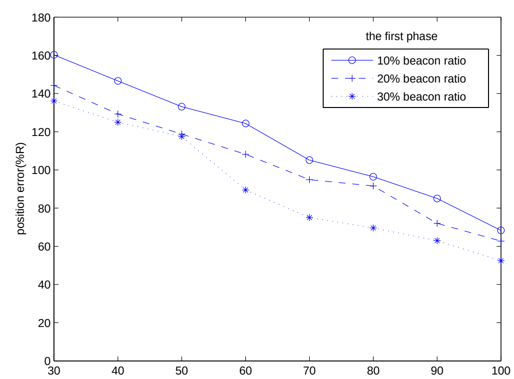

Fig. 10. The relationship between number of nodes and position error for 10%, 20%, and 30% beacon ratio in the first phase.

V. Conclusion

This paper presents a weighted multilateration positioning method that contains two phases. The first phase adopts the DV-hop method to compute the average one-hop distance and the shortest path of hop count. Then, we apply the classical multilateration to estimate their initial positions. In the second phase, non-beacon nodes exchange information with neighbors and apply the weighted multilateration method to refine their positions repeatedly. The proposed weighted multilateration positioning method only needs fewer beacons to perform the positioning function. From the simulation results, the position error of the proposed method is better than that of the RPA when the connectivity is less than 24. The limitation of the proposed method is that if the un-localized node has less than

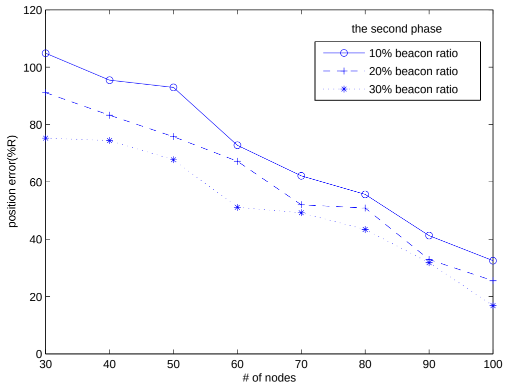

Fig. 11. The relationship between number of nodes and position error for 10%, 20%, and 30% beacon ratio in the second phase.

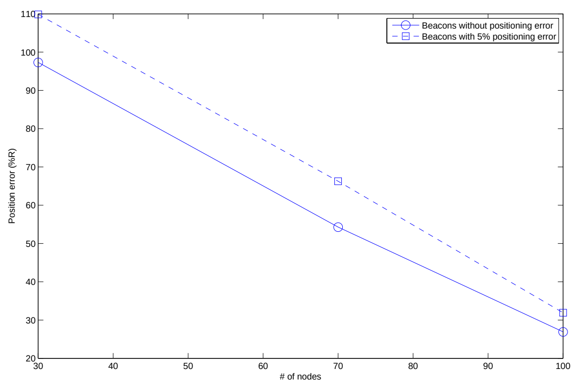

Fig. 12. The relationship between number of nodes and position error for 20% beacon ratio that each beacon node has 5% positioning error.

three neighbors, it can not be located. We will try to solve this for two ways in the future work. One may use the processing delay trick to wait for enough location information. If a un-localized node can be prior processed and successfully turn into located node, it can provide its location information to its neighbors. The other may use the n-hop message passing to obtain more location information that was came from n-hop nodes.

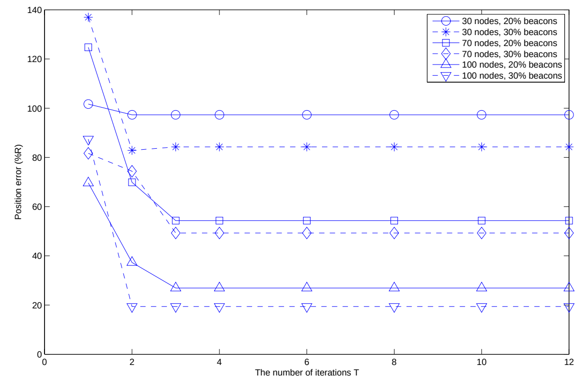

Fig. 13. The relationship between number of iterations and position error for 20% and 30% beacon ratio in the second phase.

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Yu-He Gau received the B.S. degree in Computer Science and Information Engineering from National Central University, Taiwan, in 2003 and M.S. degree in Computer and Information Science from National Chiao Tung University, Taiwan, in 2005. His research interests include wireless networks, mobile computing and wireless Internet.

Hung-Chi Chu received the B.S. and M.S. degrees in Computer Science and Engineering from Tatung University, Taiwan, in 1995 and 1997, respectively. Science 2001, he has been working toward the Ph.D. degree in Computer and Information Science at National Chiao Tung University, Taiwan. His research interests include wireless networks, mobile computing and artificial intelligence.

Rong-Hong Jan received the B.S. and M.S. degrees in Industrial Engineering, and the Ph.D. degree in Computer Science from National Tsing Hua University, Taiwan, in 1979, 1983, and 1987, respectively. He joined the Department of Computer and Information Science, National Chiao Tung University, in 1987, where he is currently a Professor. During 1991-1992, he was a Visiting Associate Professor in the Department of Computer Science, University of Maryland, College Park, MD. His research interests include wireless networks, mobile computing, distributed systems, network reliability, and operations research.