Maximizing Capacity with Power Control under Physical Interference Model in Duplex Mode

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***Abstract*—This paper addresses the joint selection and power assignment of a largest set of given links which can communicate successfully at the same time under the physical interference model in the duplex (i.e. bidirectional) mode. For the special setting in which all nodes have unlimited maximum transmission power, Halldorsson and Mitra [5] developed an approximation algorithm with a huge constant approximation bound. For the general setting in which all nodes have bounded maximum transmission power, the existence of constant approximation algorithm remains open. In this paper, we resolve this open problem by developing an approximation algorithm which not only works for the general setting of bounded maximum transmission power, but also has a much smaller constant approximation bound.**

***Index Terms*—Link scheduling, physical interference, approximation algorithms.**

I. Introduction

A wireless network is specified, in its most general format, by a triple $(V, A, \mathcal{I})$, where V is the set of networking nodes, A is the set of communication links among V, and $\mathcal{I}$ is the collection of sets of links in A which can transmit successfully at the same time. Each set in $\mathcal{I}$ is referred to as an *independent set*. The maximum size of the independent sets is called the *independence number* of the network, and is denoted by α. The problem of finding a maximum independent subset of a give set $B \subseteq A$ is referred to as **Maximum Independent Set of Links** (**MISL**). It has essential algorithmic applications in many wireless link scheduling problems including the following five ones:

- **Maximum Weighted Independent Set of Links** (**MWISL**): Given a link weight function $d \in \mathbb{R}_+^A$, find a set $I \in \mathcal{I}$ with maximum total weight $\sum_{a \in I} d(a)$.
- **Shortest Link Schedule** (**SLS**): Given a subset B of A, partition B into the smallest number of independent sets.
- **Shortest Fractional Link Schedule** (**SFLS**): A *fractional link schedule* is a set
$$\Pi = \{(I_j, l_j) \in \mathcal{I} \times \mathbb{R}_+ : 1 \leq j \leq k\};$$
its length is $\sum_{j=1}^{k} l_j$; the link load function supported by Π is the function $c_\Pi \in \mathbb{R}_+^A$ given by
$$c_\Pi(a) = \sum_{j=1}^{k} l_j |I_j \cap \{a\}|, \forall a \in A.$$
Given a link demand function $d \in \mathbb{R}_+^A$, the problem **SFLS** seeks a shortest fractional link schedule Π such that $d = c_\Pi$.
- **Maximum Multiflow** (**MMF**): Given a set of end-to-end communication requests specified by source-destination pairs, find a fractional link schedule Π of length at most one such that the maximum multiflow subject to the link capacity function c_Π is maximized.
- **Maximum Concurrent Multiflow** (**MCMF**): Given a set of end-to-end communication requests specified by source-destination pairs together with their demands, find a fractional link schedule Π of length at most one such that the maximum concurrent multiflow subject to the link capacity function c_Π is maximized.

It was recently shown in [8], [11] that if **MISL** has a polynomial μ-approximation algorithm, then both **SLS** and **SFLS** have a polynomial $(1 + \mu \ln \alpha)$-approximation algorithm, and all of the three problems **MWISL**, **MMF**, and **MCMF** all have a polynomial $e(1 + \ln \alpha)\mu$-approximation algorithm, where $e \approx 2.718$ is the natural base. Therefore, the problem **MISL** plays a vital role in the design and analysis of link scheduling algorithms in wireless networks.

This paper studies **MISL** in wireless networks under the physical interference model in the bidirectional mode. Under the physical interference model, all nodes in V lie in a metric space such as the Euclidean 3-dimentional space and the Euclidean plane. The distance between any pair of nodes u and v is denoted by $\ell(u, v)$. The signal strength attenuates with a path loss factor $\eta r^{-\kappa}$, where r is the distance from the transmitter, κ is *path-loss exponent* (a constant between 2 and 5 depending on the wireless environment), and η is the *reference loss factor*. In the bidirectional mode, the communications between a pair of nodes are bidirectional

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(see figure 1), and when two node-disjoint links a and b transmit at the same time, the interference of a toward b is $p(a) \cdot \eta\ell(a,b)^{-\kappa}$, where $\ell(a,b)$ is the shortest distance between the two endpoints of a and the two endpoints of b. The signal quality perceived by a receiver is measured by the *signal to interference and noise ratio* (*SINR*), which is the quotient between the power of the wanted signal and the total power of unwanted signals (i.e., interferences) and the ambient noise ξ. In order to correctly interpret the wanted signal, the SINR must exceed certain threshold σ. Thus, for a link a to commutate successfully even without any interference, its transmission power should exceed

$$p_0(a) = \frac{\sigma\xi}{\eta}\ell(a)^{\kappa},$$

where $\ell(a)$ is the length of a. Suppose that all nodes have maximum transmission power P. Then, the largest possible set of communication links, denoted by A, consists of all possible pairs a of nodes satisfying that $p_0(a) < P$. For a specific power assignment p, the set of communication links, denoted by A_p, consists of all possible pairs a of nodes satisfying that

$$p_0(a) < p(a) \leq P.$$

Clearly, $A_p \subseteq A$. In the duplex (i.e., bidirectional) mode of the physical interference model, the communications between a pair of nodes are bidirectional (see figure 1). For a specific power assignment p and any pair of distinct links a and b in A_p, when they transmit at the same time the interference of a toward b is $p(a) \cdot \eta\ell(a,b)^{-\kappa}$, where $\ell(a,b)$ is the shortest distance between the two endpoints of a and the two endpoints of b. Let $\mathcal{I}_p$ denote the collection of all subsets of A_p which can communicate successfully at the same time with the power assignment p.

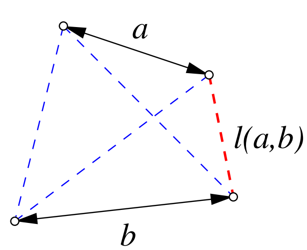

Fig. 1. The duplex mode: the communications between a pair of nodes are bidirectional, and the interference distance $\ell(a,b)$ from a link a to a link b is the shortest distance between the two endpoints of a and the two endpoints of b.

There are two variants of the **MISL** under the physical interference model in the bidirectional mode, with or without power control. In the variant without power control, a power assignment p is pre-specified and the network instance of **MISL** is simply $(V, A_p, \mathcal{I}_p)$. If $p(a)$ is a constant for all $a \in A$, then it is said be *uniform*; in particular, if $p(a) = P$ for all $a \in A$, p is referred to as the *canonical* uniform power assignment. If $p(a)$ is proportional to $\sqrt{p_0(a)}$, then it is said be *mean*; in particular, if $p(a) = \sqrt{p_0(a)P}$ for all $a \in A$, p is referred to as the *canonical* mean power assignment. If $p(a)$ is proportional to $p_0(a)$, then it is said be *linear*. A power assignment p is said to be *monotone* if $p(a)$ is non-decreasing with $\ell(a)$, to be *sub-linear* if $p(a)\ell(a)^{-\kappa}$ is non-increasing with $\ell(a)$, and to be *sub-mean* if $p(a)\ell(a)^{-\kappa}(p(a) - p_0(a))$ is non-increasing with $\ell(a)$. Clearly, all of uniform power assignments, mean power assignments, and linear power assignments are monotone and sub-linear; both uniform power assignments and mean power assignments are sub-mean. In general, any monotone and sub-mean power assignment must be also sub-linear. For **MISL** with a fixed monotone and sub-linear power assignment under the physical interference model in the bidirectional mode, a $(960 \cdot 3^{\kappa})$-approximation algorithm was developed in Halldórsson and Mitra [5].

In the variant with power control, the power assignment is part of the output rather than an input. The network instance is $(V, A, \mathcal{I})$ where $\mathcal{I}$ is the union of $\mathcal{I}_p$ for all possible power assignments p. For any $B \subseteq A$, **MISL** with power control seeks a power assignment p and a subset $I \in \mathcal{I}_p$ of B such that $|I|$ is maximized. In [5], Halldórsson and Mitra assumed *unlimited* maximum transmission power and developed an approximation algorithm with a huge constant approximation bound. The assumption on unlimited maximum transmission power is essential in their algorithm as it would technically avoid the major technical obstacle due to the ambient noise. However, such strong assumption is not only impractical, but also effectively results in the *single-hop* wireless network in which every pair of nodes can directly with each other, as opposed to the targeted multihop wireless network.

The main purpose of this paper is to develop a constant-approximation algorithm for **MISL** with power control in which all nodes have *bounded* maximum transmission power. For this purpose, we discover the suboptimality of canonical uniform power assignment and/or canonical mean power assignment. We also develop polynomial approximation algorithms for **MISL** with a fixed monotone and sub-linear power assignment, whose approximation bounds are significantly smaller than that obtained in [5]. These intermediate results are exploited to develop polynomial approximation algorithms for **MISL** with power control, which not only work for the general case of *bounded* maximum transmission power, but also have much smaller approximation bounds. Our better approximation bounds also immediately lead to improved approximations for other wireless link scheduling problems including **SLS**, **SFLS**, **MWISL**, **MMF**, and **MCMF**.

The remaining of this paper is organized as follows. Section II reviews the works related to **MISL** under physical interference model. Section III introduces the notion of independent sets in weighted digraphs and presents a simple algorithm to extract an independent set from some averagely independent set. Such algorithm and its performance are heavily used in both design and analysis of our later approximation algorithms for **MISL**. Section IV explores the rich nature of physical interference in the bidirectional mode, which will be exploited by our approximation algorithms to achieve better performance. Section V presents a general constant-approximation algorithm for **MISL** without power control. Section VI gives a constant-approximation algorithm for **MISL** with power control. Finally, Section VII concludes this paper.

II. Related Works

Most of the existing works studied **MISL** under the physical interference model in the unidirectional mode. Even in the Euclidean plane, both **MISL** with uniform power assignment [4] and **MISL** with power control [1] are NP-hard. For **MISL** with uniform power assignment in the Euclidean plane, Wan et al. [9] developed a constant-approximation algorithm, and Halldórsson and Wattenhofer [6] claimed a constant approximation algorithm but without proof and explicit approximation bound. For **MISL** with any monotone and sub-linear power assignment in arbitrary metric space, Halldórsson and Mitra [5] gave a constant-approximation algorithm. **MISL** with power control has been studied in [11], [7], [10]. In [11], Wan et al. assumed that a bounded set S of possible values of transmission power of all nodes and obtained an $O(\beta)$-approximation algorithm, where β is the power diversity of S, defined to be smallest integer k such that there exists a partition of S into k subsets in each of which any two elements differ by a factor of at most two. In [7], Kesselheim assumed unlimited maximum transmission power and developed a constant-approximation algorithm in the fading metric and a logarithmic approximation algorithm in the general metric. In the practical setting of bounded maximum transmission power, Wan et al. [10] developed a constant-approximation algorithm in the fading metric and a logarithmic approximation algorithm in the general metric.

For **MISL** under the physical interference model in the bidirectional mode which is a subject of this paper, the best-known results were obtained by Halldórsson and Mitra [5] very recently. For the variant without power control, they developed a $(960 \cdot 3^{\kappa})$-approximation algorithm with a huge constant approximation bound $960 \cdot 3^{\kappa}$. For the special case of the variant with power control in which all nodes have unlimited maximum transmission power, they developed an $(61440 \cdot 3^{\kappa})$-approximation algorithm. The assumption on unlimited maximum transmission power is essential in their algorithm as it would technically avoid the major technical obstacle due to the ambient noise. However, such strong assumption is not only impractical, but also effectively results in the *single-hop* wireless network in which every pair of nodes can directly with each other, as opposed to the targeted multihop wireless network. For the general case of the variant with power control in which all nodes have bounded maximum transmission power, the algorithm developed by Wan et al. [11] originally for the unidirectional mode can also be applied to the bidirectional mode to produce an $O(\beta)$-approximation algorithm, where β is the power diversity of the range of the possible transmission power values.

The other link scheduling problems including **SLS**, **SFLS**, **MWISL**, **MMF**, and **MCMF** have polynomial algorithmic reductions to **MISL** at the expense of increasing the approximation ratio by a factor $O(\ln \alpha)$ [8], [11]. These reductions are valid for any interference model, including the physical interference model in either the bidirectional mode or the unidirectional mode. In [6], Halldórsson and Wattenhofer claimed a constant-approximation for **SLS** with uniform power assignment in Euclidean plane. But their algorithm is wrong, and the claim has been retracted by the authors recently. Other weaker results on these problems have been obtained in [2] among others.

III. Independent Sets of Weighted Digraphs

Let G be an arc-weighted digraph on W in which each arc e has a positive weight $c(e)$. The total weight of the arcs in G is denoted by $c(G)$. The subgraph of G induced by a subset U of W is denoted by $G[U]$. For each node $v \in V$, the weighted in-degree (respectively, weighted out-degree) of v is the total weight of the arcs entering (respectively, leaving) v, and the weighted degree of v is the sum of its weighted in-degree and weighted out-degree. A subset U of W is said to be *ϕ-independent* for some $\phi > 0$ if the maximum weighted in-degree of $G[U]$ is less than ϕ, and to be *averagely ϕ-independent* for some $\phi > 0$ if its average weighted in-degree of $G[U]$ is less than ϕ.

In this section, we describe a simple algorithm **GreedyPruning** which extracts a ϕ_1-independent set U of G for some $\phi_1 > 0$ from an averagely ϕ_2-independent set S of G for some $\phi_2 \geq \phi_1$. Let k be the integer closest to $\frac{1}{2} + \frac{|S|-1}{2\phi_2/\phi_1}$ with tie broken by choosing the larger one. The algorithm runs in two phases.

- Phase 1 successively removes a vertex with the heaviest *weighted degree* among the remaining vertices until the total weight of the subgraph induced by the remaining vertices is less than $\frac{k(k-1)}{|S|-1}\phi_2$.
- Phase 2 successively removes a remaining vertex with the heaviest *weighted in-degree* until all remaining vertices are ϕ_1-independent.

The next theorem presents a lower bound on the size of the output by the algorithm **GreedyPruning**.

Theorem 3.1: The output of **GreedyPruning** has size greater than $\frac{|S|-1}{4\phi_2/\phi_1} + \frac{1}{2}$.

Proof: Let $s = |S|$ and $\phi = \phi_2/\phi_1$. The theorem holds trivially if $s - 1 \leq 2\phi$. So, we assume that $s - 1 > 2\phi$

subsequently. Then, both k and $s-k$ are at least 2 as

$$k \geq \frac{1}{2} + \frac{s-1}{2\phi} - \frac{1}{2} = \frac{s-1}{2\phi} > 1,$$

and

$$\begin{aligned} s-k &\geq s - \left(\frac{1}{2} + \frac{s-1}{2\phi}\right) - \frac{1}{2} \\ &= \left(1 - \frac{1}{2\phi}\right)(s-1) \\ &\geq \frac{s-1}{2\phi} \\ &> 1. \end{aligned}$$

We further claim that

$$k\left(1 - \frac{k-1}{s-1}\phi\right) \geq \frac{s-1}{4\phi} + \frac{1}{2}.$$

Indeed, let

$$t = \frac{1}{2} + \frac{s-1}{2\phi} - k.$$

Then, $|t| \leq \frac{1}{2}$ and $k = \frac{1}{2} + \frac{s-1}{2\phi} - t$. So,

$$\begin{aligned} &\frac{k\phi}{s-1}\left(1 - \frac{k-1}{s-1}\phi\right) \\ &= \frac{\frac{1}{2} + \frac{s-1}{2\phi} - t}{s-1}\phi\left(1 + \frac{\phi}{s-1} - \frac{\frac{1}{2} + \frac{s-1}{2\phi} - t}{s-1}\phi\right) \\ &= \left(\frac{1}{2}\left(1 + \frac{\phi}{s-1}\right) - \frac{t\phi}{s-1}\right) \\ &\quad \cdot \left(\frac{1}{2}\left(1 + \frac{\phi}{s-1}\right) + \frac{t\phi}{s-1}\right) \\ &= \frac{1}{4}\left(1 + \frac{\phi}{s-1}\right)^2 - \left(\frac{t\phi}{s-1}\right)^2 \\ &= \frac{1}{4}\left(1 + \frac{\phi}{s-1}\right)^2 - |t|^2\left(\frac{\phi}{s-1}\right)^2 \\ &\geq \frac{1}{4}\left(1 + \frac{\phi}{s-1}\right)^2 - \frac{1}{4}\left(\frac{\phi}{s-1}\right)^2 \\ &= \frac{1}{4}\left(1 + \frac{2\phi}{s-1}\right). \end{aligned}$$

Therefore,

$$\begin{aligned} k\left(1 - \frac{k-1}{s-1}\phi\right) &\geq \frac{s-1}{\phi} \cdot \frac{1}{4}\left(1 + \frac{2\phi}{s-1}\right) \\ &= \frac{s-1}{4\phi} + \frac{1}{2}. \end{aligned}$$

Thus, our claim holds.

Let m be the number of vertices at the end of phase 1. We prove $m \geq k$ by contradiction. Assume to the contrary that $m \leq k-1$. Then the number of iterations in phase 1 is

$$s - m \geq s - k + 1 \geq 3.$$

Let $U_0 = W$ and U_i be the set of vertices at the end of the i-th iteration of phase 1 for $1 \leq i \leq s-m$. Then, $|U_i| = s-i$ for each $0 \leq i \leq s-m$, and

$$c(G[U_{s-m}]) < \frac{k(k-1)}{s-1}\phi_2 \leq c(G[U_{s-m-1}])$$

For each $1 \leq i \leq s-m$, since the total weighted degree of $G[U_{i-1}]$ is $2c(G[U_{i-1}])$, the maximum weighted degree of $G[U_{i-1}]$ is at least $2c(G[U_{i-1}]) / |U_{i-1}|$. Hence,

$$\frac{c(G[U_i])}{c(G[U_{i-1}])} \leq 1 - \frac{2}{|U_{i-1}|} = \frac{s-i-1}{s-i+1}.$$

Thus,

$$\begin{aligned} \frac{c(G[U_{s-m-1}])}{c(G[W])} &= \prod_{i=1}^{s-m-1} \frac{c(G[U_i])}{c(G[U_{i-1}])} \\ &\leq \prod_{i=1}^{s-m-1} \frac{s-i-1}{s-i+1} \\ &= \frac{m(m+1)}{s(s-1)} \\ &\leq \frac{k(k-1)}{s(s-1)}. \end{aligned}$$

So,

$$\begin{aligned} c(G[U_{s-m-1}]) &\leq \frac{k(k-1)}{s(s-1)} c(G[W]) \\ &< \frac{k(k-1)}{s(s-1)} s\phi_2 \\ &= \frac{k(k-1)}{s-1}\phi_2, \end{aligned}$$

which is a contradiction. Therefore, $m \geq k$.

Let l be the number of vertices in U_{s-m}–the set of vertices at the end of phase 1–whose weighted in-degree is at least ϕ_1. Then,

$$l \leq \frac{c(G[U_{s-m}])}{\phi_1} < \frac{k(k-1)}{s-1}\frac{\phi_2}{\phi_1} = \frac{k(k-1)}{s-1}\phi.$$

Hence, at the end of phase 2,

$$\begin{aligned} |U| &\geq m - l \\ &> m - \frac{k(k-1)}{s-1}\phi \\ &\geq k - \frac{k(k-1)}{s-1}\phi \\ &= k\left(1 - \frac{k-1}{s-1}\phi\right) \\ &\geq \frac{s-1}{4\phi} + \frac{1}{2}. \end{aligned}$$

This completes the proof of the theorem. ■

Using the probabilistic method, Fanghänel et al. [3] gave a non-constructive proof of a weaker result (Proposition 3 in [3]) reinterpreted as follows: If S is ϕ_2-independent set of G for some $\phi_2 > 0$, then S contains a ϕ_1-independent set of size at least $\frac{|S|}{8\phi_2/\phi_1}$ for any $0 < \phi_1 \leq \phi_2$. They also claimed that by applying the de-randomization technique of pairwise independence their randomized existence proof can be converted to a deterministic algorithm for computing a ϕ_1-independent set of size at least $\frac{|S|}{9\phi_2/\phi_1}$. In contrast, our algorithm **GreedyPruning** is not only much simpler, both conceptually and in implementation, but also achieves a performance more than twice better.

IV. Relative Interference Graph

Consider a power assignment p on A. The *relative interference* (also termed *affectance* in [5]) of a link $a \in A_p$ toward another link $b \in A_p \setminus \{a\}$, denoted by $RI_p(a,b)$, is defined as follows: If a and b share a common node, then $RI_p(a,b) = \infty$; otherwise,

$$RI_p(a,b) = \sigma \frac{p(a)\,\ell(a,b)^{-\kappa}}{(p(b) - p_0(b))\,\ell(b)^{-\kappa}}.$$

The *inward relative interference* of a link $a \in A_p$ from a subset $B \subseteq A_p \setminus \{a\}$ is defined to be

$$RI_p(B,a) = \sum_{b \in B} RI_p(b,a);$$

The *outward relative interference* of a link $a \in A_p$ to a subset $B \subseteq A_p \setminus \{a\}$ is defined to be

$$RI_p(a,B) = \sum_{b \in B} RI_p(a,b).$$

It's easy to verify for any subset I of A_p, $I \in \mathcal{I}_p$ (in other words, I is independent under the power assignment p) if and only for each $a \in I$,

$$RI_p(I \setminus \{a\}, a) < 1.$$

Motivated by this observation, we define the *relative interference digraph* of the power assignment p, to be the complete weighted digraph G_p on A_p in which the weight of each arc (a,b) is $RI_p(a,b)$. Then, $\mathcal{I}_p$ is exactly the collection of all independent sets of G_p. In this section, we explore more useful properties of relative inteferences.

In general, a subset B of A_p is said to be *ϕ-independent* for some $\phi > 0$ if it is a ϕ-independent set of G_p, and to be *averagely ϕ-independent* for some $\phi > 0$ if it is an averagely ϕ-independent set of G_p. Suppose that B is an averagely ϕ-independent for some $\phi \geq 1$. Then, we can apply the algorithm **GreedyPruning** to extract an independent set I from B, and by Theorem 3.1,

$$|I| > \frac{|B| - 1}{4\phi} + \frac{1}{2} > \frac{|B|}{4\phi}.$$

The following two lemmas on the relative interferences among at most three links can be verified in a straightforward manner, and therefore we skip their proofs.

Lemma 4.1: Consider a power assignment p and two links $a, b \in A_p$ with $\ell(a) \leq \ell(b)$.

1) If p is sub-mean, then $RI_p(a,b) \geq RI_p(b,a)$.
2) If p is linear, then $RI_p(a,b) \leq RI_p(b,a)$.

Lemma 4.2: Consider a monotone and sub-linear power assignment p. Then for any three links $a_1, a_2, b \in A_p$ satisfying that $\ell(a_1) \leq \ell(a_2)$ and b is disjoint from a_1 and a_2,

$$\max\left\{\frac{RI_p(a_1,b)}{RI_p(a_2,b)}, \frac{RI_p(b,a_1)}{RI_p(b,a_2)}\right\} \leq \left(\frac{\ell(a_2,b)}{\ell(a_1,b)}\right)^{\kappa}.$$

Consider a link $a \in A$ and a subset $B \subseteq A$. A subset S of B is said to be a t-guarding subset of B to a for some $t > 0$ if for each link $b \in B \setminus S$,

$$\ell(a,b) \geq t \min_{a' \in S} \ell(a',b).$$

The lemma below is a consequence of Lemma 4.2.

Lemma 4.3: Consider a monotone and sub-linear power assignment p, a link $a \in A_p$, and a set B of disjoint links in A_p. Let S be a t-guarding subset of B to a for some $t > 0$. Then for any $b \in B \setminus S$, b is disjoint from a, and if each link in S is not shorter than a then

$$RI_p(a,b) \leq t^{-\kappa} RI_p(S,b),$$
$$RI_p(b,a) \leq t^{-\kappa} RI_p(b,S).$$

Proof: Consider any link $b \in B \setminus S$. Let $a' \in S$ be such that $\ell(a,b) \geq t\ell(a',b)$. Since a' and b are disjoint, $\ell(a',b) > 0$. Thus, $\ell(a,b) > 0$ and hence b is disjoint from a. Now suppose that each link in S is not shorter than a. Then, $\ell(a) \leq \ell(a')$. By Lemma 4.2 we have

$$\frac{RI_p(a,b)}{RI_p(a',b)} \leq \left(\frac{\ell(a',b)}{\ell(a,b)}\right)^{\kappa} \leq t^{-\kappa},$$
$$\frac{RI_p(b,a)}{RI_p(b,a')} \leq \left(\frac{\ell(b,a')}{\ell(b,a)}\right)^{\kappa} \leq t^{-\kappa}.$$

Thus,

$$RI_p(a,b) \leq t^{-\kappa} RI_p(a',b) \leq t^{-\kappa} RI_p(S,b),$$
$$RI_p(b,a) \leq t^{-\kappa} RI_p(b,a') \leq t^{-\kappa} RI_p(b,S).$$

So, the lemma holds. ∎

The next lemma presents upper bounds on the size of a minimum t-guarding subset of some $B \subseteq A$ to a link $a \in A$ in various settings.

Lemma 4.4: Consider a link $a \in A$ and a subset B of of links in A.

1) There exists a $1/2$-guarding subset S of B to a such that $|S| \leq 2$.
2) If the metric space is the Euclidean plane (respectively, 3-dimensional Euclidean space), then there exists a 1-guarding subset S of B to a such that $|S| \leq 10$ (respectively, 24).

Proof: If $a \in B$, all the three parts hold trivially by choosing $S = \{a\}$. So, we assume that $a \notin B$. Let u and v be the two endpoints of a.

(1) We give a constructive proof. Initially, S is set to empty and B' is set to B. Repeat the following iterations while B' is non-empty: Let a' be a link in B' closest to a, add a to S, and remove a' and all other links b in B' satisfying that $\ell(a,b) \geq \ell(a',b)/2$. We claim that $|S| \leq 2$. Assume to the contrary that $|S| > 2$. By symmetry, we assume that at least two links in S are no farther from u than from v, and let a_1, a_2 be these links in the order of selection. For each $1 \leq j \leq 2$, let w_j denote the endpoint of a_j closer to u_i. Then,

$$\ell(a,a_1) = \ell(u,w_1) \leq \ell(u,w_2) = \ell(a,a_2) < \ell(a_1,a_2)/2.$$

However,

$$\begin{aligned}\ell(a_1,a_2) &\leq \ell(w_1,w_2) \\ &\leq \ell(u,w_1) + \ell(u,w_2) \\ &= \ell(a,a_1) + \ell(a,a_2) \\ &\leq 2\ell(a,a_2)\end{aligned}$$

which implies that

$$\ell(a,a_2) \geq \ell(a_1,a_2)/2,$$

which is a contradiction.

(2). We give a constructive proof. Initially, S is set to empty and B' is set to B. Repeat the following iterations while B' is non-empty: Let a' be a link in B' closest to a, add a to S, and remove a' and all other links b in B' satisfying that $\ell(a,b) \geq \ell(a',b)$. We claim that $|S| \leq 10$ (respectively, 24). Assume to the contrary that $|S| > 10$ (respectively, 24). By symmetry, we assume that at least half of the links in S are no farther from u than from v, and let $a_1, a_2, \cdots, a_k$ be these links in the order of selection. Then $k \geq 6$ (respectively, 13). For each $1 \leq j \leq k$, let w_j denote the endpoint of a_j closer to u_i. Then, for any $1 \leq i < j \leq k$,

$$\ell(u,w_i) \leq \ell(u,w_j) = \ell(a,a_j) < \ell(a_i,a_j) \leq \ell(w_i,w_j).$$

Therefore, $u \neq w_i$, for otherwise $\ell(u,w_j) = \ell(w_i,w_j)$ which contradicts to that $\ell(u,w_j) < \ell(w_i,w_j)$. This implies that none of w_i for $1 \leq i \leq k$ coincides with u. In addition, the angle separation between w_i and w_j at u is more than 60^o. Therefore, k is no more the number of rays emanating from u whose mutual angle separations are all greater than 60^o. In the Euclidean plane (respectively, 3-dimensional Euclidean space), this number is at most 5 (respectively, the 3-dimesnional kissing number 12), which is a contraction. Thus, our claim holds. ∎

V. MISL WITH FIXED POWER ASSIGNMENT

In this section, we present a general approximation for **MISL** with a fixed monotone and sub-linear power assignment p. Suppose B is a subset of A_p. Our approximation algorithm produces an independent set of B by taking a relaxation approach. Fix a parameter $\phi \geq 1$ whose value will be determined later on. Our algorithm **RelaxIS**(ϕ) runs in two phases:

- Phase 1 constructs an averagely ϕ-independent subset W of B. The construction of W is the first-fit selection in the increasing order of link lengths. Initially, W is empty. At each iteration, the shortest link is removed from the candidate list and it is added to W if it together with W form an averagely ϕ-independent set.
- Phase 2 applies the algorithm **GreedyPruning** to extract an independent set I from W.

The next theorem provides the proper choices of the parameter ϕ and the corresponding approximation bounds when the deployment region is an Euclidean plane.

Theorem 5.1: In an Euclidean plane, let $\phi = 2$, $\mu = 80$ if p is sub-mean; and $\phi = 2 + \sqrt{2}$, $\mu = 80\left(\frac{3}{2} + \sqrt{2}\right)$ otherwise. Then, **RelaxIS**(ϕ) has approximation bound μ.

Proof: Let W be the set of links at the end of the first phase of **RelaxIS**(ϕ), and I^* be a maximum independent set in B with the power assignment p. We define a subset O of I^* as follows.

- If p is sub-mean, O is simply I^* itself.
- If p is not sub-mean, O is the set of links in I^* whose weighted out-degree in $G_p[I^*]$ is at most $\sqrt{2}+1$. Then,

$$|O| \geq \left(1 - \frac{1}{\sqrt{2}+1}\right)|I^*| = \frac{|I^*|}{1 + \frac{\sqrt{2}}{2}},$$

  and the maximum weighted in-degree (respectively, out-degree) of $G_p[O]$ is less than one (respectively, $\sqrt{2}+1$).

It is sufficient to show that $|O| \leq 10|W|$. Indeed, if p is sub-mean, then by Theorem 3.1

$$|I| \geq \frac{|W|}{4\phi} \geq \frac{|O|}{40\phi} = \frac{|I^*|}{40\phi} = \frac{|I^*|}{\mu};$$

if p is not sub-mean, then by Theorem 3.1

$$|I| \geq \frac{|W|}{4\phi} \geq \frac{|O|}{40\phi} \geq \frac{|I^*|}{40\left(1 + \frac{\sqrt{2}}{2}\right)\phi} = \frac{|I^*|}{\mu}.$$

Next, we prove $|O| \leq 10|W|$ by contradiction.

Assume to the contrary that $|O| > 10|W|$. For any two links a and b, we use $a \prec b$ to indicate that a precedes b in the increasing order of link lengths adopted by the first phase of **RelaxIS**(ϕ), and use $a \preceq b$ to indicate either $a = b$ or $a \prec b$. We shall assign exclusive "guards" from O to each link in W. Initialize O' to O. For each $a \in W$ in the order of addition to W, $S(a)$ be a smallest 1-guarding subset of

$$\{b \in O' : a \preceq b\}$$

to a and remove $S(a)$ from O'. By Lemma 4.4, $|S(a)| \leq 10$ for each $a \in W$. Since $|O| > 10|W|$, none of $S(a)$ for $a \in W$ and $O \setminus (\cup_{a \in W} S(a)$ is empty. Choose an arbitrary link $b \in O \setminus (\cup_{a \in W} S(a)$. By Lemma 4.3, for each $a \in W$, b is disjoint from a and

$$RI_p(a,b) \le RI_p(S(a), b),$$
$$RI_p(b,a) \le RI_p(b, S(a)),$$

Consequently, $b \notin W$. Let

$$W' = \{a \in W : a \prec b\}.$$

We claim that $RI_p(W', b) + RI_p(b, W') < \phi$. We prove the claim in two cases.

Case 1: p is sub-mean. Then

$$\begin{aligned} RI_p(W', b) &= \textstyle\sum_{a \in W'} RI_p(a,b) \\ &\le \textstyle\sum_{a \in W'} RI_p(S(a), b) \\ &< 1, \end{aligned}$$

and by Lemma 4.1,

$$RI_p(W', b) + RI_p(b, W') \le 2RI_p(W', b) < 2 = \phi.$$

Case 2: p is not sub-mean, then

$$\begin{aligned} RI_p(W', b) &= \textstyle\sum_{a \in W'} RI_p(a,b) \\ &\le \textstyle\sum_{a \in W'} RI_p(S(a), b) \\ &< 1, \end{aligned}$$

and

$$\begin{aligned} RI_p(b, W') &= \textstyle\sum_{a \in W'} RI_p(b,a) \\ &\le \textstyle\sum_{a \in W'} RI_p(b, S(a)) \\ &< \sqrt{2} + 1. \end{aligned}$$

Therefore,

$$RI_p(W', b) + RI_p(b, W') < 2 + \sqrt{2} = \phi.$$

The above claim implies that $W' \cup \{b\}$ would form an averagely ϕ-independent set and hence b should have been added to W by the algorithm **RelaxIS**(ϕ), which is a contradiction. ∎

The next two theorem provide the proper choices of the parameter ϕ and the corresponding approximation bounds when the deployment region is a 3-dimensional Euclidean space or an arbitrary metric space.

Theorem 5.2: In 3-dimensional Euclidean space, let $\phi = 2$, $\mu = 192$ if p is sub-mean; and $\phi = 2+\sqrt{2}$, $\mu = 192\left(\frac{3}{2} + \sqrt{2}\right)$ otherwise. Then, **RelaxIS**(ϕ) has approximation bound μ.

Theorem 5.3: In an arbitrary metric space, let $\phi = 2^{\kappa+1}$, $\mu = 2^{\kappa+4}$ if p is sub-mean; and $\phi = 2^{\kappa}\left(2+\sqrt{2}\right)$, $\mu = 2^{\kappa+4}\left(\frac{3}{2} + \sqrt{2}\right)$ otherwise. Then, **RelaxIS**(ϕ) has approximation bound μ.

The proofs of the above two theorems is similar to that of Theorem 5.1 and are omitted.

The approximation bounds μ in the above three theorems are significantly smaller the approximation bound $960 \cdot 3^{\kappa}$ of the algorithm by Halldórsson and Mitra [5]. Indeed, consider the setting with $\kappa = 3$ and a sub-mean power assignment p. The approximation bound in [5] is

$$960 \cdot 3^{\kappa} = 960 \cdot 3^3 = 25920;$$

while our approximation bound is 80 in Euclidean plane (with $\phi = 2$) and $2^7 = 128$ in an arbitrary metric space (with $\phi = 16$), both of which are more than 200 times smaller than $25,920$. Another nice feature of our algorithm is that in Euclidean plane or 3-dimensional Euclidean space its approximation bound does not depend on the value of the path-loss exponent κ.

Finally, we remark that while the best choice of the parameter ϕ for the **RelaxIS**(ϕ) is unknown, we may guess it algorithmically. Specifically, we can choose a set Φ of a polynomial number of candidates. For each $\phi \in \Phi$, we run the algorithm to **RelaxIS**(ϕ) to select an independent set. The largest one among those selected independents is the final output.

VI. MISL with Power Control

In this section, we present an approximation **IS/PC** for **MISL** with power control. A cornerstone to the algorithm **IS/PC** is the suboptimality of the canonical uniform power assignment and/or the canonical mean power assignment, which will be established in this section. A polynomial approximation algorithm $\mathcal{A}$ for **MISL** with monotone and sub-mean power assignment, such as the one developed in the previous section, is exploited by the algorithm **IS/PC** as a subroutine. Denote by q_1 (respectively, q_2) the canonical uniform (respectively, mean) power assignment, and let B be a subset of A. The algorithm **IS/PC** produces an independent set of B with power control in the following steps.

- Step 1: Apply the algorithm $\mathcal{A}$ to B and q_1 to select a subset $I_1 \in \mathcal{I}_{q_1}$ of B.
- Step 2: Apply the algorithm $\mathcal{A}$ to B and $q_{1=2}$ to select a subset $I_2 \in \mathcal{I}_{q_2}$ of B.
- Step 3: If $|I_1| \ge |I_2|$, output I_1 and q_1;otherwise, output I_2 and q_2.

The next theorem given an approximation bound of the algorithm **IS/PC**.

Theorem 6.1: Suppose that the algorithm $\mathcal{A}$ has an approximation bound μ. Then, algorithm **IS/PC** has an approximation bound $8g^2\mu$, where g is the golden ratio $\frac{1+\sqrt{5}}{2}$.

The remaining of this section is devoted the proof of Theorem 6.1. It's easy to verify that $A_{q_1} = A_{q_2} = A$. In addition, canonical uniform power assignment and canonical mean power assignment have the following properties.

Lemma 6.2: Consider a power assignment p and a set $I \in \mathcal{I}_p$.

1) If $\max_{a \in I} p(a) \leq t \min_{a \in I} p(a)$ for some $t > 1$, then I contains a subset $I' \in \mathcal{I}_{q_1}$ such that $|I'| \geq \frac{1}{4t}|I|$.
2) If $\max_{a \in I} q_2(a) \leq P/t$ for some $t > 1$, then then I contains a subset $I' \in \mathcal{I}_{q_2}$ such that $|I'| \geq \frac{1-1/t}{4}|I|$.

Proof: (1). Since $p(b) \leq P$, we have

$$\begin{aligned}&\frac{RI_{q_1}(a,b)}{RI_p(a,b)}\\ &= \frac{q_1(a)}{p(a)}\frac{p(b)-p_0(b)}{q_1(b)-p_0(b)}\\ &= \frac{P}{p(a)}\frac{p(b)-p_0(b)}{P-p_0(b)}\\ &\leq \frac{P}{p(a)}\frac{p(b)}{P} = \frac{p(b)}{p(a)} \leq t.\end{aligned}$$

Therefore, I is t-independent with the power assignment q. By Theorem 3.1, the first part of the lemma holds.

(2). Consider any two distinct links a and b in I. We claim that

$$\frac{RI_{q_2}(a,b)+RI_{q_2}(b,a)}{RI_p(a,b)+RI_p(b,a)} < \frac{1}{1-1/t}.$$

Indeed,

$$\begin{aligned}&\frac{RI_{q_2}(a,b)}{RI_{q_2}(b,a)}\\ &= \frac{q_2(a)(q_2(a)-p_0(a))}{q_2(b)(q_2(b)-p_0(b))}\left(\frac{\ell(b)}{\ell(a)}\right)^{\kappa}\\ &= \frac{q_2(a)(q_2(a)-p_0(a))}{q_2(b)(q_2(b)-p_0(b))}\left(\frac{q_2(b)}{q_2(a)}\right)^{2}\\ &= \frac{1-\frac{p_0(a)}{q_2(a)}}{1-\frac{p_0(b)}{q_2(b)}} = \frac{1-\frac{q_2(a)}{P}}{1-\frac{q_2(b)}{P}}.\end{aligned}$$

Since

$$\begin{aligned}&\frac{RI_p(a,b)\,RI_p(b,a)}{RI_{q_2}(a,b)\,RI_{q_2}(b,a)}\\ &= \frac{p(a)\,p(b)\,(q_2(a)-p_0(a))\,(q_2(b)-p_0(b))}{q_2(a)\,q_2(b)\,(p(a)-p_0(a))\,(p(b)-p_0(b))}\\ &> \frac{(q_2(a)-p_0(a))\,(q_2(b)-p_0(b))}{q_2(a)\,q_2(b)}\\ &= \frac{(P-q_2(a))\,(P-q_2(b))}{P^2}\\ &= \left(1-\frac{q_2(a)}{P}\right)\left(1-\frac{q_2(b)}{P}\right),\end{aligned}$$

we have

$$\begin{aligned}&RI_p(a,b)\,RI_p(b,a)\\ &> \left(1-\frac{q_2(a)}{P}\right)\left(1-\frac{q_2(b)}{P}\right)RI_{q_2}(a,b)\,RI_{q_2}(b,a)\\ &= \left[\left(1-\frac{q_2(a)}{P}\right)RI_{q_2}(b,a)\right]^2.\end{aligned}$$

which further implies

$$\begin{aligned}&\left(1-\frac{q_2(a)}{P}\right)RI_{q_2}(b,a)\\ &< \sqrt{RI_{q_2}(a,b)\,RI_{q_2}(b,a)}\\ &\leq \frac{RI_p(a,b)+RI_p(b,a)}{2}.\end{aligned}$$

Thus,

$$\begin{aligned}&RI_{q_2}(a,b)+RI_{q_2}(b,a)\\ &= \left(1+\frac{1-\frac{q_2(a)}{P}}{1-\frac{q_2(b)}{P}}\right)RI_{q_2}(b,a)\\ &= \left(\frac{1}{1-\frac{q_2(a)}{P}}+\frac{1}{1-\frac{q_2(b)}{P}}\right)\left(1-\frac{q_2(a)}{P}\right)RI_{q_2}(b,a)\\ &\leq \left(\frac{1}{1-\frac{q_2(a)}{P}}+\frac{1}{1-\frac{q_2(b)}{P}}\right)\frac{RI_p(a,b)+RI_p(b,a)}{2}.\end{aligned}$$

So,

$$\begin{aligned}&\frac{RI_{q_2}(a,b)+RI_{q_2}(b,a)}{RI_p(a,b)+RI_p(b,a)}\\ &< \frac{1}{2}\left(\frac{1}{1-\frac{q_2(a)}{P}}+\frac{1}{1-\frac{q_2(b)}{P}}\right)\\ &\leq \frac{1}{1-1/t}.\end{aligned}$$

Therefore, our claim holds.

The above claim implies that the total relative interference of I with the power assignment q_2 is less than $\frac{1}{1-1/t}$ times the total relative interference of I with the power assignment p,and hence is less than $\frac{1}{1-1/t}|I|$. So, I is averagely $\frac{1}{1-1/t}$-independent with the power assignment q_2. By Theorem 3.1, the second part of the lemma holds. ■

We classify the links in A into long links and short links, and treat them separately. A link $a \in A$ is called a *long* link if $p_0(a) \geq P/g^2$, and a *short* link otherwise. Using the equality that $1 - 1/g = 1/g^2$, we obtain the following corollary from Lemma 6.2.

Corollary 6.3: If a set I of long (respectively, short) links is independent with power control, then at least $|I|/(4g^2)$ links in I are independent with canonical uniform (respectively, mean) power assignment.

Now, we are ready to prove Theorem 6.1. Let I_1 and I_2 be the two sets output at Step 1 and Step 2 respectively, and let O be a maximum IS of B with power control and p be a power assignment satisfying that $O \in \mathcal{I}_p$. Denote by O_1 (respectively, O_2) the set of long (respectively, short) links in I. By Corollary 6.3, at least $|O_1|/(4g^2)$ links in O_1 are independent with canonical uniform power assignment, and at least $|O_2|/(4g^2)$ links in O_2 are independent with canonical mean power assignment. Since the algorithm $\mathcal{A}$ has approximation bound μ, we have

$$|I_1| \geq \frac{|O_1| / \left(4g^2\right)}{\mu} = \frac{|O_1|}{4g^2\mu},$$
$$|I_2| \geq \frac{|O_2| / \left(4g^2\right)}{\mu} = \frac{|O_2|}{4g^2\mu}.$$

Thus,

$$\max\{|I_1|, |I_2|\} \geq \frac{\max\{|O_1|, |O_2|\}}{4g^2\mu} \geq \frac{|O|}{8g^2\mu}.$$

So, Theorem 6.1 holds.

Theorem 6.1 implies that if the algorithm **IS/PC** adopts the algorithm given in the previous section as the algorithm $\mathcal{A}$, then its approximation ratio is at most $10 \cdot (8g)^2$ if the metric space is an Euclidean plane, and $2^{\kappa+7}g^2$ otherwise.

For the "short" instances of B, the algorithm **IS/PC** may have an even smaller approximation bound. Let $R = \left(\frac{\eta P}{\sigma\xi}\right)^{1/\kappa}$. Then, all links in A have length less than R. Thus, it is proper to measure the shortness of B by the parameter

$$\lambda = \frac{\max_{a\in B} \ell(a)}{R}.$$

Let O be a maximum IS of B with power control. Since

$$\frac{P}{\max_{a\in B} q(a)} = \lambda^{-\kappa/2},$$

at least $\frac{1-\lambda^{\kappa/2}}{4}|O|$ links in O are independent with canonical mean power assignment by Lemma 6.2. Thus,

$$|I_2| \geq \frac{1-\lambda^{\kappa/2}}{4\mu}|O|,$$

which implies the algorithm **IS/PC** has an approximation bound of at most $\frac{4\mu}{1-\lambda^{\kappa/2}}$. Note that when $\lambda = \left(\sqrt{5}-1\right)^{-2/\kappa}$,

$$\frac{4}{1-\lambda^{\kappa/2}} = 8g^2.$$

Therefore, when $\lambda < \left(\sqrt{5}-1\right)^{-2/\kappa}$, I_2 alone is already an $8g^2\mu$-approximate solution. For another example, if $\kappa = 4$ and $\lambda \leq \frac{\sqrt{2}}{2} \approx 0.707$, I_2 alone is a 8μ-approximate solution.

Under the assumption of unlimited maximum transmission power, Halldórsson and Mitra [5] gave an approximation algorithm with approximation bound $61440 \cdot 3^\kappa$. Under this assumption, the shortness λ of any set B is zero trivially, and hence I_2 alone is a 4μ-approximate solution. By adopting the algorithm presented in the previous section, the approximation bound 4μ is significantly smaller than $61440 \cdot 3^\kappa$. For example, consider the setting with $\kappa = 3$. The approximation bound in [5] is

$$61440 \cdot 3^\kappa = 61440 \cdot 3^3 = 1,658,880;$$

while our approximation bound is 320 in Euclidean plane and 512 in an arbitrary metric space, both of which are more than $3,000$ times smaller than $1,658,880$.

VII. Conclusion

In this paper, we have developed an approximation algorithm for **MISL** with power control under the physical interference model in the bidirectional mode. Our approximation algorithm is the first constant-approximation algorithm which does not assume unbounded maximum transmission power. All prior works on **MISL** with power control in either bidirectional mode or unidirectional mode have to assume unbounded maximum transmission power for the purpose of avoiding the technical obstacle due to the ambient noise. Furthermore, our approximation algorithm significantly outperforms the one presented in [5] which assumed unbounded maximum transmission power. Both the design and analysis of our approximation algorithm are built upon the rich nature of the physical interference in bidirectional mode discovered in this paper. These properties are of independent interest and are expected to find applications in other wireless scheduling problems.

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