Low Energy Fault Tolerant Bounded-Hop Broadcast in Wireless Networks

Hanan Shpungin*
Computer Sciences Dept.
Ben-Gurion University of the Negev
Beer-Sheva, Israel
shpungin@cs.bgu.ac.il

Michael Segal†
Communication Systems Engineering Dept.
Ben-Gurion University of the Negev
Beer-Sheva, Israel
segal@cse.bgu.ac.il

ABSTRACT
This paper studies asymmetric power assignments in wireless ad-hoc networks. The temporary and unfixed physical topology of wireless ad-hoc network is determined by the distribution of the wireless nodes as well as the transmission power (range) assignment of each node. We consider the problem of bounded-hop broadcast under k-fault resilience criterion for linear and planar layout of nodes. The topology which results from our power assignment allows a broadcast operation from a wireless node r to any other node in at most h hops and is k-fault resistant.

We develop simple approximation algorithms for the two cases and obtain the following approximation ratios: linear case – O(k); planar case – we first prove a factor of O(k3), which is later decreased to O(k2) by a finer analysis. Finally we show a trivial power assignment with a cost O(h) times the optimum. To the best of our knowledge these are the first non-trivial results for this problem.

Categories and Subject Descriptors
C.2.1 [Computer-Communication Networks]: Network Architecture and Design—Wireless communications; G.2.2 [Discrete Mathematics]: Graph Theory—Network Problems

General Terms
Algorithms, Design, Reliability, Theory

Keywords

1. INTRODUCTION
A wireless ad-hoc network consists of several transceivers (stations), communicating by radio. Each transceiver t is assigned a transmission power p(t), which gives it some transmission range, denoted by r_t. This is customary to assume that the minimal transmission power required to transmit to a distance d is d^\alpha, where the distance-power gradient \alpha is usually taken to be in the interval [2, 4] (see [31]). Thus, a transceiver t receives transmissions from s if \( p(s) \geq d(s,t)^\alpha \), where \( d(s,t) \) is the Euclidean distance between s and t. The transmission possibilities resulting from a power assignment induce a communication graph. Research efforts have focused on finding power assignments, for which the induced communication graph satisfies a certain topology property, while minimizing the total cost. Broadcasting is one of possible topologies. For a special station r, called the root, we want to establish a transmission graph where there is route from r to every other node in the network. In some cases it is essential to minimize the number of hops from the root transceiver to other nodes in the network. This is called bounded-hop broadcasting.

This paper is organized as follows. In the rest of this section we present the model, previous work and briefly describe our results. In Section 3 we address the linear case of the problem. Then in Section 3 we deal with the planar layout of nodes and prove various approximation factors for the problem.

1.1 The Model
We are given a set \( T \) of n transceivers \( t_1, t_2, \ldots, t_n \), positioned in \( \mathbb{R}^d \), \( d \geq 1 \). We define the cost of an undirected graph \( G(T) = (T, \mathcal{E}) \) by

\[
C_G = \sum_{(s,t) \in \mathcal{E}} c(s,t).
\]

With edge costs being \( c(s,t) = d(s,t)^\alpha \) for every \( (s,t) \in \mathcal{E} \). When each transceiver is assigned a transmission power \( p(t) = r_t^\gamma \), an ad-hoc network is created. A power assignment for \( T \) is a vector of transmission powers \( \{p(t) \mid t \in T\} \) and is denoted by \( A(T) \) (usually abbreviated to \( A \)). The resulting (directed) communication graph is denoted by \( H_A = (T, \mathcal{E}_A) \), where \( \mathcal{E}_A \) is the set of directed edges resulting from the power assignment \( A(T) \):

\[
\mathcal{E}_A = \{(t,s) \mid p(t) \geq d(t,s)^\alpha \}.
\]

That is, there is a directed edge from \( t \) to \( s \) if \( t \) has sufficient transmission power to reach \( s \). In this paper we refer to
transceivers as nodes. The cost of the power assignment is defined as the sum of all transmission powers:
\[ C_A = \sum_{t \in T} p(t). \]

Throughout this paper we address the linear layout of nodes as well. The transceivers are positioned on a single line in increasing order from right to left (see Figure 1.a). Note that in case of the linear layout if there is a path from \( t_i \) to \( t_j \), where \( i < j \) then there is a path from \( t_i \) to any node \( t_l \), where \( i < l < j \).

For some power assignment \( A \) and a root node \( r \), we say that a communication graph \( H_A \) is a broadcast graph rooted at \( r \) if for any other node \( t \in T \) there is a path from \( r \) to \( t \) in \( H_A \). In the case that the maximal number of hops from \( r \) to any node \( t \) is limited by some constant \( h \), we say that \( H_A \) is an \( h \)-bounded-hop broadcast graph rooted at \( r \). In this paper we demand that in case of the linear layout if there is a path from \( t_i \) to any node \( t_l \), where \( i < l < j \), then there is a path from \( t_i \) to \( t_j \), where \( i < j \).

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Note that \( k \)-fault resistance definition above presumes that there are \( k \)-node disjoint paths from \( r \) to any node \( t \), which is not directly reached by \( r \). Since if some node \( u \) is reached by the root, any \( k-1 \) node removals will not effect the transmission from \( r \) to \( t \). On the other hand, if some node \( v \) is node directly reached by \( r \), then there must exist \( k \) node disjoint paths from \( r \) to \( v \), otherwise a removal of \( k-1 \) nodes might result in transmission failure from \( r \) to \( v \). In this work we assume \( \alpha = 2 \).

1.2 Previous Work

Topology control in wireless networks is a relatively new field of interest. Nevertheless a wide area of problems has already been studied. Most of the problems are aimed at computing a low energy power assignment that meets global topological constraints. Kouris et al. [28] introduced the MinRange(SC) problem, which is the \( k \)-strong connectivity problem for \( k = 1 \). They proved it to be NP-Hard for the three dimensional Euclidean space for any value of \( \alpha \). The same paper provided a 2-approximation algorithm for the planar case and an exact \( O(n^4) \) time algorithm for the one dimensional case. In the planar case, the NP-Hardness of the problem for every \( \alpha \) has been proved in [19] and a simple 1.5-approximation algorithm for the case \( \alpha = 1 \) has been provided in [5]. Ambuhl et al. [4] presented some algorithms for the weighted power assignment, solving it optimally for the broadcast, multi-source broadcast and strong connectivity problems for the linear case (they achieved the same running time for the strong connectivity problem as in [28]). They also presented some approximation algorithms for the multi-dimensional case. An excellent survey covering many variations of the problem is given in [17].

A natural generalization of the strong connectivity requirement is \( k \)-strong connectivity. These networks also provide multi-path redundancy for load balancing or transmission fault tolerance. As power-optimal strong connectivity is NP-Hard, so is power-optimal \( k \)-strong connectivity. Two versions of the problem arise: symmetric and asymmetric. In the symmetric version for any two nodes \( t, s \in T \), \( p(t) \geq d(t, s)^\alpha \Leftrightarrow p(s) \geq d(s, t)^\alpha \), that is a node \( t \) can reach node \( s \) if and only if \( s \) can reach node \( t \), we can also refer to it as an undirected model. The asymmetric version allows directed links between two nodes. Krumke et al. [29] argued that the asymmetric version is harder than the symmetric version. A first non trivial result for planar asymmetric \( k \)-strong connectivity was presented by Shpungin and Segal in [33]. They derived an approximation factor of \( O(k^2) \) for the planar case and some results for the linear case. Carmi et al. [14] improved the approximation ratio to \( O(k) \). Another possible connectivity property is \( k \)-edge connectivity, which implies that the removal of any \( k \) edges results in a disconnected graph. In [13], Calinescu and Wan presented various aspects of symmetric/asymmetric \( k \)-strong connectivity and \( k \)-edge connectivity. They first proved NP-Hardness of the symmetric two-edge and two-node strong connectivity and then provided a 4-approximation algorithm for both symmetric and asymmetric strong biconnectivity \( (k = 2) \) and a \( 2k \)-approximation for both symmetric and asymmetric \( k \)-edge strong connectivity. Hajiaghayi et al. [25] give two algorithms for symmetric \( k \)-strong connectivity, with \( O(k \log k) \) and \( O(k) \)-approximation factors and also some distributed approximation algorithms for \( k = 2 \) and \( k = 3 \) in geometric graphs. Jia et al. in [27] present various approximation factors (depending on \( k \)) for the symmetric \( k \)-strong connectivity, such as \( 3k \)-approximation algorithm for any \( k \geq 3 \) and \( 6 \)-approximation for \( k = 3 \). Segal and Shpungin [32] extend static algorithms for \( k \)-connectivity to support dynamic node insert/delete operations. Additional results can be found in [1, 9, 12, 16, 20, 26, 30, 34].

Wieselthier et al. in [36, 37] were the first to study the broadcast problem in wireless ad-hoc networks for the 2-dimensional case and when \( \alpha = 2 \). In this work, the performances of three heuristics, namely the minimum spanning tree (MST), the shortest path tree (SPT) and the broadcasting incremental power (BIP) have been experimentally compared (one to each other) on the random uniform model without providing theoretical results. The approach taken in [36, 37] is to build a source rooted spanning tree by adjusting transmit powers of nodes, followed by a sweep operation to remove redundant transmissions. Wan et al. in [35] present the first analytical results for this problem by exploring geometric structures of Euclidean MSTs. In particular, they prove that the approximation ratio of MST is between 6 and 12, for BIP it is between \( \frac{4}{3} \) and 12 and for SPT it is at least \( \frac{2}{3} \), where \( n \) is the number of receiving nodes given that there are no obstacles in the network and that the fixed energy cost for electronics is negligible. Cagalj et al. [10] give a proof of NP-Hardness of the minimum-energy broadcast problem in a Euclidean space. Many researchers provided
analytic results of the minimum-energy broadcast algorithm
based on computing an MST. In [2] Ambuhl et al. proves
an approximation factor of 6, which matches the lower bound
previously known for this algorithm. Flammini et al. [23] es-
tablish improved approximation results on the performance
of BIP. Cartigny et al. in [15] develop localized algorithms
for minimum-energy broadcasting. Segal and Shpungin [32]
develop a general framework for k-fault tolerance in various
topology problems and provide an approximation bound of
$O(k^3)$ for the k-broadcast problem. Additional references
and results may be found in [2, 6, 8].

We can also add an additional constraint parameter to the
problem, the bounded diameter $D$ of the induced communica-
tion graph. For the linear case node disposition, Kirousis
et al. [28] develop an optimal power assignment algorithm
in $O(n^2)$ time. In the Euclidean case, [21] obtains con-
stant ratio algorithms for the bounded-hop strong connect-
ivity for well spread instances. Beier et al. [7] discuss the
problem of finding a bounded-hop path between pairs of
nodes with minimized power consumption. They find an
optimal path in $O(hn \log n)$ time. In [11] the authors obtain
$O((\log n) \log n)$ bicriteria approximation algorithms for
the bounded-hop broadcast, bounded-hop connectivity and
bounded-hop symmetric connectivity problems. In their
output there are at most $h \log n$ hops with $\log n$ times the
optimal cost for $h$ hops. In [3] the authors present an exact
algorithm for solving the 2-hop broadcast problem with a
running time of $O(n^3)$ as well as a PTAS with a running
time of $O(n^\mu)$ where $\mu = O((\log^2 h) \log h)$. Funke and Laue [24]
provide a PTAS for the h-broadcast algorithm in time linear
in $n$. Additional results for bounded range assignments can
be found in [18, 22].

1.3 Our Contribution
We study the problem of $h$-bounded broadcasting in con-
junction with k-fault resistance (MEkBHB). We first pro-
vide a $O(k)$ approximation algorithm for the linear layout of
nodes. For the planar case we develop an approximation al-
gorithm, with provable approximation ratio of $O(k^3)$, which
is later decreased to $O(k^2)$ by a fine analysis. Finally we
show a trivial power assignment with a cost $O(h)$ times the
optimum. All our algorithms run in low polynomial time.

2. LINEAR BOUNDED-HOP BROADCAST
Given $n$ nodes positioned on a single line, we assume they are
located in an increasing order from left to right (see Figure
1.a). We start by solving a special case of the MEkBHB
problem for the linear layout of nodes, when the root node
is $r = t_1$, that is from the leftmost node.

2.1 Broadcast from $t_1$
Let $\Delta = \max_{1 \leq i \leq n} d(t_i, t_{i+1})$ be maximal distance between
two adjacent nodes and $D = d(t_1, t_n)$ be the distance from
the leftmost to rightmost nodes. We assume $\Delta k \leq \frac{D}{2}$, that
is if we divide the line into sections of length $D/h$ then each
section will contain at least $k$ nodes. The next Lemma gives
the lower bound for the cost of the optimal power assignment
$A^*$ under these settings.

**Lemma 2.1.** $C_{A^*} \geq \frac{D^2 k^2}{(k+h-1)^2}$.

![Figure 1: Linear $k$-fault resistant bounded-hop broadcast from $t_1$ for $k = 3$, $h = 4$.](image)

**Proof.** Let $t_i$ be the rightmost node reached by $t_1$ in $A^*$.
Nodes that are not directly reached by $t_1$ must be accessible
by $k$-node disjoint paths from it. In case of the linear layout,
if there is a path from $u$ to $v$, then there is a path from $u$ to
any node in between. Therefore in $A^*$ there must be $k$ node
disjoint traversals to $t_n$ originating nodes $t_i, t_{i+k}, \ldots, t_{i+j}$,
since paths originating in earlier nodes will not be optimal.
Let $x = d(t_1, t_i)$. The distance covered by these paths is
$D - x$ (see Figure 1.b). Note that $C_{A^*}$ is minimized if all
these paths start at $t_i$ and evenly divide the distance $D - x$
into $h - 1$ hops. Given that the root node $t_1$ is assigned the
range $x$ we can bound the optimal power assignment by:

$$C_{A^*} \geq x^2 + \left(\frac{D - x}{h - 1}\right)^2 k(h - 1).$$

Next we analyze the function $f(x) = x^2 + \left(\frac{D - x}{h - 1}\right)^2 k(h - 1)$ to find the lower bound for the cost of the optimal power assignment. For $x = \frac{Dk}{h+k-1}$, the value of function $f(x)$ is minimized. As a result,

$$f\left(\frac{Dk}{h+k-1}\right) = \frac{D^2 k^2}{(k+h-1)^2} + \left(\frac{D - \frac{Dk}{h+k-1}}{h - 1}\right)^2 k(h - 1)$$

$$= \frac{D^2 k^2}{(k+h-1)^2} + \frac{D^2 k(h-1)}{(k+h-1)^2}$$

$$= \frac{D^2 k^2}{(k+h-1)^2} = \frac{D^2 k}{k+h-1},$$

which completes our proof. \hfill \blacksquare

Next we describe our algorithm. We divide the line into $h$
blocks of length $D/h$ each. Note that due to our assumption,
there are at least $k$ nodes in each block. Let $t_{ij}$ be the
rightmost node in the \(j\)th block. We assign power as follows. The root is assigned \(p(t_1) = d(t_1, t_{i_1})^2\) to reach \(t_{i_1}\). In blocks 1 to \(h-2\), each of the \(k\) rightmost nodes is assigned with enough power to reach the \(k\) rightmost nodes of the next block. In block \(h-1\), all \(k\) rightmost nodes are assigned with enough power to reach \(t_n\) and nodes in the last block are assigned zero power (see Figure 1.c). Formally for \(1 \leq l \leq k\),

\[ p(t_{i_l} - k + l) = d(t_{i_l} - k + l, t_{i_{j_l} - k + l})^2, \quad \text{for } 1 \leq j < h - 1 \]

\[ p(t_{i_h} - k + l) = d(t_{i_h} - k + l, t_n)^2. \]

Let \(L_h^k\) be the resulting power assignment. Easy to see that the the induced (directed) communication graph \(H_{L_h^k}\) is \(k\)-fault resistant \(h\)-bounded-hop broadcast graph rooted at \(t_1\). It is sufficient to show the existence of \(k\)-node disjoint paths from \(t_1\) to \(t_n\). These paths can be described as,

\[ y_l = (t_1, t_{i_1} - k + l, t_{i_2} - k + l, \ldots, t_{i_h} - k + l, t_n), \quad 1 \leq l \leq k. \]

**Lemma 2.2.** \(C_{L_h^k} \leq \frac{D^2}{h} (4k(h - 1) + 1).\)

**Proof.** The root is assigned a transmission range of at most \(D/h\). A total of \(k(h-1)\) nodes are assigned a transmission range of at most \(D/h + k\Delta\) each. Recall our assumption \(\Delta k \leq \frac{D}{h}\). Therefore,

\[ C_{L_h^k} \leq \left( \frac{D^2}{h^2} + k(h-1) \left( \frac{D}{h} + k\Delta \right) \right)^2 \leq \left( \frac{D^2}{h^2} + k(h-1) \left( \frac{2D}{h} \right) \right)^2 = \frac{D^2}{h} (4k(h - 1) + 1). \]

This completes our proof. \(\blacksquare\)

Finally we can easily derive our main Theorem.

**Theorem 2.3.** \(C_{L_h^k} \in O(k)C_{A^*}.\)

**Proof.** According to Lemmas 2.1 and 2.2 we have \(C_{A^*} \geq \frac{D^2k}{k(h-1)}\) and \(C_{L_h^k} \leq \frac{D^2}{h} (4k(h - 1) + 1).\) Also, since \(k \geq 1\) and \(h \geq 2\) then \(k + h - 1 \leq kh\). Therefore,

\[ \frac{C_{L_h^k}}{C_{A^*}} \leq \frac{D^2}{h} (4k(h - 1) + 1) \left( \frac{k + h - 1}{k} \right) = \frac{4k(h - 1) + 1}{h} \leq 4k + 1. \]

We conclude \(C_{L_h^k} \in O(k)C_{A^*}.\) \(\blacksquare\)

### 2.2 Broadcast from any \(t_i\)

The generalization of our algorithm to broadcast from any node \(t_i\), \(1 \leq i \leq n\) is very simple. We use the described algorithm to obtain two power assignments \(A_L\) and \(A_R\); the former is a power assignment for \(k\)-\(h\)-broadcast from \(t_1\) to nodes to the left of it – namely \(t_1, \ldots, t_{i-1}\); the latter is for nodes to the right of it – namely \(t_{i+1}, \ldots, t_n\). Next we combine the two power assignments. Let \(p_L(t)\) and \(p_R(t)\) be the power assigned to \(t\) in \(A_L\) and \(A_R\) respectively. We define the power assignment \(L(t)\) as follows, \(p(t) = \max\{p_L(t), p_R(t)\}\).

Clearly the induced communication graph \(H_A\) is \(k\)-fault resistant and \(h\)-bounded-hop broadcast from \(t_1\). Let \(A^*_t\) be the optimal power assignment for \(k\)-\(h\)-broadcast from \(t_i\).

Since solving the problem for a subset of adjacent nodes will produce a cheaper solution, we can use the bound in Theorem 2.3 and conclude \(C_{A_L} \in O(k)C_{A^*_t}\) and \(C_{A_R} \in O(k)C_{A^*_t}\). Therefore \(C_{L(i)^*} \in O(k)C_{A^*_t}.\) Easy to see that after sorting the nodes the running time of the algorithm is \(O(n \log n)\).

### 3. Planar Bounded-hop Broadcast

The general idea for a power assignment which induces a \(k\)-fault resistant bounded-hop broadcast graph in the plane \(s\) to first obtain a bounded-hop broadcast graph and then make it \(k\)-fault resistant. In [24] the authors present a PTAS algorithm for the MEkBHB problem with fault resistance parameter \(k = 1\). We use this construction as a basis for \(k\)-fault resistant bounded-hop broadcast. We first explain the technique used for \(k\)-fault resistant strong connectivity suggested in [14], that we will use later to obtain \(k\)-fault resistant broadcast tree.

#### 3.1 Planar \(k\)-strong connectivity

Let \(T\) be a set of \(n\) points in the plane (representing \(n\) transceivers). For each node \(t \in T\), let \(N_t \subseteq T\) be a set of \(k\)-closest nodes to \(t\), and put \(r_t^* = \max_{i' \in N_t} d(t, t')\). We now describe the power assignment algorithm. Compute a minimum spanning tree MST of the Euclidean graph induced by \(T\). Assign to each node \(t \in T\) a power \(p(t) = (r_t^*)^2\). As a result, each node can reach its \(k\)-closest neighbours. Denote this initial range assignment by \(A^*\). For each edge \(e = (t, s)\) of MST, increase the power of the nodes in \(N_t \cup N_s\) (if necessary), such that each node \(t' \in N_t\) can reach all nodes in \(N_s\), and vice versa. Let \(A_k\) denote the resulting power assignment.

The idea is rather simple, we want to construct \(k\)-node disjoint paths along the edges of the MST. Think about each \(N_t\) as large intersections containing \(k\) intersection points, and that there are \(k\) symmetric links between \(N_t\) and \(N_s\) iff \((t, s)\) is an edge in the MST. The range assignment of each node \(t\) should be at least \(r_t^*\) (otherwise \(k\)-strong connectivity is impossible), and in addition sufficient enough to create the intersections mentioned above. The following Lemma can be easily proved (see Figure 2).

**Lemma 3.1.** Given two nodes \(t, s \in T\), let \(r_t^*, r_s^*\) be the range node \(t' \in N_t\) has to be assigned in order to reach any node in \(N_s\) including \(s\). Then, \(r_t^* \leq r_s^* + d(t, s) + r_s^*\).

We provide a proof sketch for the following Theorem from [14]. Let \(A_s^*\) be an optimal power assignment for the \(k\)-connectivity problem.

**Theorem 3.2.** (Carmi et al. [14]). \(C_{A_k} \in O(k)C_{A^*_t}\)

**Proof Sketch.** Note that the range assignment of each node \(t \in A_k\) must satisfy the following two conditions; (a) it must reach at least \(k\) other nodes (b) it must satisfy the demands of other nodes, for which it is one of their \(k\)-closest neighbors (i.e. all those nodes \(s\) so that \(t \in N_s\)).
For an edge $e = (t, s)$ in the MST we denote $r^e_{t, s} = r^e_{t', s'}$. Therefore each node $t'$ is assigned a transmission range which is the maximum between $r^e_{t, s}$ and its obligation to some node $t$. Where $t' \in N_t$ and there is an edge $e \in \text{MST}$ so that $r^e_{t, s} > r^*_{t'}$ (see Figure 2). Therefore,

$$
C_A = \sum_{t' \in \mathcal{T}} p(t') \leq \sum_{t' \in \mathcal{T}} \max \left\{ r^e_{t', s} , \max_{e \in \text{MST}} \{ r^e_{t, s} \} \right\} .
$$

From Lemma 3.1 and the fact that geometrical MST has a bounded degree of 6,

$$
C_A \leq O(k) \left( \sum_{t' \in \mathcal{T}} (r^e_{t'})^2 + \sum_{e \in \text{MST}} |e|^2 \right).
$$

According to [28] $C_{\text{MST}} \leq C_{A^1} \leq C_{A^k}$. We can conclude,

$$
C_A = O(k) \left( C_{A^1} + C_{\text{MST}} \right) = O(k) C_{A^1}.
$$

**3.2 The algorithm**

Given a set of nodes $\mathcal{T}$ and a root node $r$, we wish to construct a power assignment $A^k_r$, so that the induced communication graph $H^k_{Ak}$ is $k$-h-hop-rooted at $r$. As before, for each node $t \in \mathcal{T}$, let $N_t \subseteq \mathcal{T}$ be a set of $k$-closest nodes to $t$, and put $r^*_{t} = \max_{t' \in N_t} d(t, t')$. Let $A_k$ be a power assignment constructed in [24] for some constant $h$, so that $H^k_{Ak}$ is a 1-h-hop broadcast graph. We are ready to describe the power assignment algorithm.

We start by constructing a directed spanning tree of $H^k_{Ak}$ by running a BFS from the root node $r$. Denote the resulting tree by $\text{BHT}'$ and by level-$i$ nodes to be the nodes at distance $i$ from the root. Clearly for each node $t \in \mathcal{T}$ there is a unique directed path of at most $h$ hops from $r$ to $t$ in the $\text{BHT}'$. Note that the power assignment $A^k_{\text{BHT}}$ required to induce this tree has a cost $C_{A^k_{\text{BHT}}}$ (see Figure 3). Call this tree BHT and by $A_{\text{BHT}}$ the power assignment required to induce this tree. Easy to see that $C^k_{A_{\text{BHT}}} \leq 2C_{A^k_{\text{BHT}}} \leq 2C_A$ by using the following observation.

**Observation 3.3.** For any $x_1, x_2, \ldots, x_m \in \mathbb{N}$ it holds

$$
\left( \sum_{i=1}^{m} x_i \right)^2 \leq m \sum_{i=1}^{m} x_i^2 .
$$

Similar to the case of strong connectivity, we would like to create $k$-node disjoint paths along the edges of $\text{BHT}$, from $r$ to any node other $t \in \mathcal{T}$. We start by assigning the root node $r$ with a power $p(r) = (r^*_r)^2$, so that it can reach its $k$-closest neighbors. Next, for each directed edge $e = (t, s)$ (from $t$ to $s$) in $\text{BHT}$ we increase the power (if required) of each node $t' \in N_t$ so it could reach all nodes in $N_s \cup \{s\}$. Let $A^k_r$ denote the resulting power assignment.

It is easy to see that the resulting (directed) communication graph $H^k_{Ak}$ is $k$-fault resistant $h$-bounded-hop broadcast rooted at $r$. That is, for any node $t \in \mathcal{T}$ there are $k$-node disjoint paths from $r$ to $t$, that "follow" the path from $r$ to $t$ in $\text{BHT}$. And each of these paths has at most $h$ hops.

**3.3 Analysis**

In order to analyze the cost of the power assignment $A^k_r$, we need to take a closer look at the power increase stage of each node. All nodes (except for the root) start with no power and it is increased if required. The power of $t' \in N_t$ is increased only to satisfy the demand of some outgoing edge $e = (t, s)$ from $t$ in $\text{BHT}$, that is to reach any node in $N_s \cup \{s\}$. Since node $t'$ can be a member in many sets of $k$-closest neighbors, it might be required to increase its power many times, but eventually its power will be dominated by some outgoing edge $e' = (t_i, t_s)$, where $t' \in N_{t_i}$. Recall that for an edge $e = (t, s)$ we denote by $r^e_{t, s}$ the range node $t'$ has to be assigned to reach $s$ and all the $k$-closest neighbors of $s$.

To simplify the notation, for any node $t_i$, denote by $e_i = (t(i), s(i))$ the edge which dominates the power assignment of $t'_i$, where $t(i) \neq s(i)$ are some nodes in $\mathcal{T}$. Note that it is possible that $t(i) = t(j)$ for $i \neq j$. The root node might not have a dominating edge since its initial power is greater than 0. However, we will assume it has one and later show that it does not influence our analysis at all.

![Figure 2: Node $t'$ is assigned a range of at most $r^*_t + d(t, s) + r^*_s$ to reach all $s' \in N_s$.](image)

![Figure 3: Depth decrease of $\text{BHT}'$.](image)
\begin{lemma}
\label{lem:approximation-ratio-bound}
C_{A^h_k} \in O(k) \left( \sum_{t \in T} (r_t^*)^2 + C_{A^h_k} \right).
\end{lemma}

\textbf{Proof.}
From Lemma \ref{lem:property-of-power-assignment} we have the following inequality: 
\begin{align*}
r_t^{(s(i), s(i))} + d(t(i), s(i)) + r_t^*(s(i)).
\end{align*}
From Observation \ref{obs:approximation-ratio-bound}, \begin{align*} 
 p(t(i)) \leq \left( (r_t^*)^2 + d(t(i), s(i))^2 + (r_t^*)^2 \right). 
\end{align*}
Let \( p(t) \) be the power node \( t \) is assigned in \( A_{BHT} \). Then \begin{align*} 
 p'(t(i)) \leq d(t(i), s(i))^2. 
\end{align*}
We can write,
\begin{align*}
C_{A^h_k} = \sum_{i=1}^{n} p(t_i) \leq 3 \sum_{i=1}^{n} \left( (r_t^*)^2 + p'(t(i)) + (r_t^*)^2 \right).
\end{align*}
For any node \( t \in T \), only for \( t_i \in N_t \) we have \( t(i) = t \). For any node \( s \in T \), let \( e_s = (t_s, s) \) be an incoming edge of \( s \) in HBT. Then only for \( t_i \in N_t \) we have \( s(i) = s \). As a result, 
\begin{align*} 
\sum_{i=1}^{n} (r_t^*)^2 \leq k \sum_{i \in T} (r_t^*)^2, \quad \sum_{i=1}^{n} (r_t^*)^2 \leq k \sum_{i \in T} (r_t^*)^2 
\end{align*}
and 
\begin{align*} 
\sum_{i=1}^{n} p'(t(i)) \leq k \sum_{i \in T} p'(t) = k C_{A_{BHT}} \leq 2k C_{A^h_k}.
\end{align*}
Therefore, \begin{align*} 
C_{A^h_k} = O(k) \left( \sum_{i \in T} (r_t^*)^2 + C_{A^h_k} \right). 
\end{align*}
We note that if the root has a dominating edge it does not affect the analysis.
\hfill \Box

Paper \cite{Papakostas2006} proves that given power assignment algorithm which produces the \( A^h_k \) assignment is PTAS and therefore it holds \( C_{A^h_k} \leq (1 + \epsilon) C_{A^h_k} \), where \( A^h_k \) and \( A^h_k \) are the optimal power assignments that induce a 1-h broadcast and \( k \)-h broadcast communication graphs respectively. In the next two Lemmas, which prove two different approximation ratios of our algorithm, the cost of \( C_{A^h_k} \) is negligible in relation to \( \sum_{i \in T} (r_t^*)^2 \).

\begin{theorem}
\label{thm:approximation-ratio-bound}
C_{A^h_k} \in O(k^3) C_{A^h_k}.
\end{theorem}

\textbf{Proof.}
Let every node \( t \in T \) be assigned a transmission range \( r_t^* \). Call this power assignment \( A_1 \). The induced communication graph holds the following property, \textit{every node has at least} \( k \) \textit{neighbors}. We also claim that \( A_1 \) is an optimal power assignment that induces a communication graph with such a property.

An approximation algorithm for a power assignment \( A^{k-1}_h \) which induces a \((k-n-1)\)-broadcast communication graph (there is no constant bound on the path length) is given in Segal and Shpunig \cite{Segal2002} and they prove that \( C_{A^{k-1}_h} \in O(k^2) C_{A^{k-1}_h} \), where \( A^{k-1}_h \) is the optimal power assignment for the \( k \)-fault resistant unbounded broadcast.

The induced communication graph \( H_{A^{k-1}_h} \) maintains the property that \textit{every node has at least} \( k \) \textit{neighbors}. Therefore we can conclude that \( C_{A^h_1} \leq C_{A^{k-1}_h} \). Easy to see that 
\begin{align*}
C_{A^{k-1}_h} \leq C_{A^h_k}, 
\end{align*}
since forcing a constant maximal number of hops increases the power assignment. As a result, 
\begin{align*}
\sum_{t \in T} (r_t^*)^2 = C_{A^h_1} \in O(k^2) C_{A^h_k}.
\end{align*}
In conjunction with Lemma \ref{lem:approximation-ratio-bound} we conclude
\begin{align*}
C_{A^h_k} \in O(k) \left( \sum_{t \in T} (r_t^*)^2 + C_{A^h_k} \right) = O(k^3) C_{A^h_k}.
\end{align*}
We will need the following two Lemmas to prove Theorem \ref{thm:approximation-ratio-bound}.

\begin{lemma}
\label{lem:approximation-ratio-bound}
ANY single point \( p \) in the plane can not be covered by more than \( 6k + 6 \) disks, so that each disk does not cover more than \( k \) centers of other disks.
\end{lemma}

\textbf{Proof.}
We divide the plane into 6 equal sectors of \( \pi/3 \) and show that at most \( k + 1 \) disks that cover \( p \) can reside in each sector so that each disk does not cover more than \( k \) centers of other disks. Suppose, by contrary, that one of the sectors contains more than \( k + 1 \) disks. In Figure 4 there is a sector \( \beta \) with 4 disk centers (\( k = 2 \)). Let \( x \) be the center of a disk in \( \beta \) at largest distance most from \( p \). Let \( D_p \) be a disk centered at \( p \) with a radius \( d(p, x) \). Clearly, all the other disk centers in the sector \( \beta \) are covered by \( D_p \) and there are at least \( k + 1 \) of them. Since we have chosen a sector \( \beta \) of \( \pi/3 \), all these disk centers belong to \( D_e \cap D_p \) and therefore are covered by \( D_e \). Contradiction.
\hfill \Box

\begin{lemma}
\label{lem:approximation-ratio-bound}
\textit{Let} \( S \) be a set of \( m > k \) \textit{nodes in the plane positioned inside disk} \( D \) \textit{of radius} \( r \). \textit{Let} \( A^* \) \textit{be the optimal power assignment so that each node can reach at least} \( k \) \textit{nodes in} \( S \). \textit{Then} \( C_{A^*} \in O(k)r^2 \).
\end{lemma}

\textbf{Proof.}
Let \( p^*(s_i) \) be the power node \( s_i \in S \) is assigned in \( A^* \). Let \( D_i \) be a disk centered at node \( s_i \) as a result of the power assignment (i.e. the disk is centered at \( s_i \) and has a radius of \( \sqrt{p^*(s_i)} \)). Since \( A^* \) is an optimal power assignment, then each disk can cover at most \( k \) centers of other nodes. Due to Lemma \ref{lem:approximation-ratio-bound} each point inside disk \( D \) can be covered by at most \( O(k) \) disks, therefore
\begin{align*}
\sum_{i=1}^{m} \text{area}(D_i) \in O(k) \text{area}(D) \Rightarrow \sum_{i=1}^{m} p^*(s_i) \in O(k)r^2.
\end{align*}
\hfill \Box
Let $C_{A_3}(Z_u)$ be the total power assigned to nodes of $Z_u$ in $A_3$. If we combine the two cases we obtain for every $u \in T$, $C_{A_3}(Z_u) \in O(k)p(u)$. Therefore we can conclude,

$$C_{A_3} = \sum_{u \in T} C_{A_3}(Z_u) + (x_r)^2 = O(k) \sum_{u \in T} p(u) = O(k)C_{H_{A_k^{n-1}}^{3}}.$$  

We claim every node in $H_{A_k}$ has at least $k$ neighbors. Easy to see that the root has $k$ neighbors. Note that $|Z_r| \geq k$, because the root necessarily has at least $k$ neighbors in $A_k^{n-1}$. For any other node $t$:

**Case 1:** If $|Z_u| < k$ then $u \neq r$ and necessarily $|R_t| \geq k$. Also $p_3(t) = p_2(t)$ and therefore node $t$ reaches all the nodes in $R_t$.

**Case 2:** If $|Z_u| \geq k$ then it is assigned to reach $k$ different nodes in $Z_u$.

We have obtained a power assignment $A_3$ so that the induced communication graph $H_{A_k}$ satisfies the property that every node has at least $k$ neighbors. As a result, $C_{A_1} \leq C_{A_3}$. We showed that $C_{A_3} \in O(k)C_{H_{A_k^{n-1}}^{3}}$. Similarly to the proof of Theorem 3.5 and in conjunction with Lemma 3.4 we conclude that

$$C_{A_k} \in O(k) \left( \sum_{r \in T} (r_i)^2 + C_{A_h} \right) = O(k^2)C_{A_k}^{h}.$$  

It takes linear time to construct $H_{A_h}$. Easy to see that our algorithm runs in a total of $O(n^2)$ time.

**A simple $O(h)$ approximation for very high fault resistance** - Instead of forming $k$-node disjoint paths from $r$ to any other node, we could simply assign the root with enough power to reach all nodes in a single hop. Clearly such a power assignment is very fault resistant since the transmission between $r$ and any other node does not rely on relay nodes.

**Lemma 3.9.** Let $t$ be the most distant node from $r$. Let $A_h$ be a power assignment where the root is assigned $p(r) = d(r,t)^2$. Then $C_{A_h} \in O(h)C_{A_h}^{h}$.

**Proof.** There is a path $y = (r = u_1, u_2, \ldots, u_{t+1} = t)$ of at most $l \leq h$ hops from $r$ to $t$ in $H_{A_h}$. Let $r_u$ be the range assigned to node $u$ in $A_h$. Clearly $p(r) \leq \left( \sum_{i=1}^{t} r_{u_i} \right)^2$ and $\sum_{u=1}^{l} r_{u_i}^2 \leq C_{A_h}$. According to Observation 3.3 we have $p(r) \leq \left( \sum_{i=1}^{t} r_{u_i} \right)^2 \leq l \left( \sum_{i=1}^{t} r_{u_i} \right)^2 \leq hC_{A_h}$. We conclude $C_{A_h} \in O(h)C_{A_h}$ and as a result $C_{A_h} \in O(h)C_{A_h}^{h}$.

**4. CONCLUSIONS**

In this paper we showed numerous results for the fault tolerant bounded hop broadcast topology problem in wireless networks. A possible interesting direction would be to improve the analysis of the approximation ratios obtained in this work or show a ratio which is $k,h$-dependent. It might be also of interest to develop a distributed approach for such a problem.
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5. REFERENCES


