Approximating Dense Cases of Covering Problems

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Abstract

We study dense instances of several covering problems. An instance of the set cover problem with \( m \) sets is dense if there is \( \epsilon > 0 \) such that any element belongs to at least \( \epsilon m \) sets. We show that the dense set cover problem can be approximated with the performance ratio \( c \log n \) for any \( c > 0 \) and it is unlikely to be NP-hard. We construct a polynomial-time approximation scheme for the dense Steiner tree problem in \( n \)-vertex graphs, i.e. for the case when each terminal is adjacent to at least \( \epsilon n \) vertices. We also study the vertex cover problem in \( \epsilon \)-dense graphs. Though this problem is shown to be still MAX-SNP-hard as in general graphs, we find a better approximation algorithm with the performance ratio \( \frac{2}{1+\epsilon} \). The superdense cases of all these problems are shown to be solvable in polynomial time.

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1 Dense Set Cover Problem

We start with the Dense Set Cover Problem. Let \( X = \{x_1, ..., x_k\} \) be a finite set and \( P = \{p_1, ..., p_m\} \subseteq 2^X \) be a family of its subsets. The Set Cover Problem (SCP) asks for a minimum size sub-family \( M \) of \( P \) such that \( X \subseteq \bigcup\{p \mid p \in M\} \).

The greedy heuristic gives \( 1 + \ln k \) approximation for SCP \([5]\). Moreover, SCP cannot be approximated to within less than \( \ln k \)-factor unless \( NP \subseteq DTIME[n^{\log \log n}] \) \([6]\).

The \( B \)-sparse SCP has a constant upper bound \( B > 1 \) on the number of sets in \( P \) which cover the same element of \( X \). The Vertex Cover Problem is a well-known representative of \( B \)-sparse SCP (\( B=2 \)). There is a simple \( B \)-approximation algorithm for this problem. From the other side, the \( B \)-sparse SCP is MAX SNP-complete.

In an \( \epsilon \)-dense SCP, any element of \( X \) belongs to at least \( \epsilon |P| \) sets for some \( \epsilon < 1 \).

We will analyze the greedy heuristic applied to \( \epsilon \)-dense SCP. This heuristic repeatedly choose a maximum size set in \( P \), remove its elements from \( X \) and all other sets in \( P \). All chosen sets form the output set cover Greedy.

**Lemma 1** The size of Greedy is at most \( \log_{1/(1-\epsilon)}(k) \).

**Proof.** At first we will show that the maximum size of a set in \( P \) is at least \( \epsilon k \). Consider a bipartite graph \( G = (P \cup X, E) \) where \( x \in X \) and \( p \in P \) are adjacent if and only if \( x \in p \). The degree of any \( x \in X \) is at least \( \epsilon m \), so the number of edges in this graph is at least \( \epsilon mk \) and, therefore, there is a set \( p \in P \) with degree at least \( \epsilon m \).

Each iteration of the greedy heuristic does not decrease density, since all elements which belong to the chosen set are removed from \( X \). So the size of \( X \) after the \( i \)th iteration is at most \( (1 - \epsilon)^i k \).

This lemma shows that the size of the optimal set cover is \( O(\log k) \). So we cannot expect that the \( \epsilon \)-dense SCP is \( NP \)-complete, since a simple \( O(m^{O(\log k)}) \)-time exhaustive search chooses the optimal solution.

**Theorem 1** Unless \( NP \subseteq DTIME[n^{\log n}] \), the \( \epsilon \)-dense SCP is not \( NP \)-complete.

Note that \( O(\log k) \) is the tight bound for the performance ratio of the greedy heuristic applied to \( \epsilon \)-dense SCP. To show this for \( \epsilon = \frac{1}{2} \), we can construct an instance of this problem with the size of optimal solution of \( O(\log k) \) and then add two sets \( A \) and \( B \) such that \( A \cup B = X \), \( A \cap B = \emptyset \). On the other hand, unlike to the general case of SCP, we may decrease the constant factor as far as we want.

**Lemma 2** For any \( \epsilon > 0 \) and \( 1 > \epsilon > 0 \), there is a \( \epsilon \ln k \)-approximation algorithm for \( \epsilon \)-dense SCP.
Proof. Indeed, let transform an instance of $\epsilon$-dense SCP to an instance of 
$(1 - (1 - \epsilon)^2)$-dense SCP in the following way. Consider a family $P^2 = \{p \cup q : p, q \in P\}$. It is easy to see that any solution for SCP with the family $P^2$ gives a solution for initial SCP. An $\epsilon$-density means that at most $(1 - \epsilon)m$ sets do not contain a given element of $X$. But then at most $(1 - \epsilon)^2m^2$ sets in $P^2$ do not contain a given element of $X$. 

Lemma 1 implies that such transformation decrease the performance ratio of the greedy algorithm twice. ◇

Theorem 1 arises the following two open problems:

**Problem 1** Can $\epsilon$-dense SCP be solved in polynomial time?

**Problem 2** Can $\epsilon$-dense SCP be approximated in polynomial time to within constant factor?

Further densification leads to polynomial solvability of SCP. The $\delta$-superdense SCP is the case of SCP where each element of $X$ is covered by at least $m - o(m^\delta)$ sets of $P$ for some $\delta < 1$.

**Theorem 2** The $\delta$-superdense SCP can be solved in polynomial time.

**Proof.** Let each element of $X$ is covered by at least $m - \gamma m^\delta = m(1 - \gamma m^{\delta - 1})$ sets of $P$ for some $\gamma < m^{1 - \delta}$. By Lemma 1 for $\epsilon = 1 - \gamma m^{\delta - 1}$, the size of optimal solution is at most

$$\log_{\gamma - 1} m^{1 - \delta} k = \frac{1}{(1 - \delta)(1 - \log_m \gamma)} \log_m k.$$ 

Thus, exhaustive search for finding an exact solution has at most $k^{(1 - \delta)\delta^{-1}}$ cases to consider. ◇

## 2 Dense Steiner Tree Problem

Consider a connected graph $G = (V, E)$ with a terminal set $S \subseteq V$. The Steiner Tree Problem (STP) asks for a minimum size tree within $G$ which spans all terminals from $S$. Further, $d(F)$ denotes the length of a graph $F$, $|S| = k$ and $|V| = n$. A well-known minimum spanning tree heuristic (MSTH) [9] finds a minimum spanning tree $M$ of a weighted complete graph $G' = (S, E', c)$, where the weight of any edge equals to the length of the shortest path between its ends in $G$. Then MSTH replaces all edges of $M$ with the corresponding paths in $G$ and extracts a tree from the subgraph obtained.

An optimal Steiner tree contains also non-terminals. Each such vertex of degree at least 3 is called a Steiner point. It is easy to see that there are at most $k - 2$ Steiner points. Using MSTH we can find an optimal Steiner tree if we add all Steiner points to the terminal set.
Remark 1 An optimal Steiner tree can be found exactly in $O(n^k)$ time.

MSTH gives 2-approximation for STP [9] and the best up-to-date polynomial-time approximation guarantee is about 1.644 [7]. From the other side, STP is known to be MAX SNP-complete [4].

In the $B$-sparse STP the degree of any vertex is bounded by a constant $B$. It is known that STP in the rectilinear metric (a sub-case of 4-sparse STP) is $NP$-complete but the question whether it is MAX SNP-hard or not is still open.

In an $\epsilon$-dense instance of STP (for some $\epsilon < 1$) any terminal has at least $\epsilon n$ neighbors outside $S$.

Note that for $\epsilon > \frac{1}{2}$, $\epsilon$-dense STP is a sub-case of Network STP with distances 1 and 2 which is still MAX SNP-complete [4]. The Rayward-Smith heuristic [8] was proposed for the latter problem in [4]. It achieves a better approximation guarantee ($\frac{3}{2}$) then MSTH which has the tight bound 2 as for the general case. MSTH also does not differ the dense and general case of STP.

If the number of terminals is small enough, i.e. $k \leq \frac{1}{\epsilon}$, then we can find an exact solution in polynomial time. Otherwise, we apply to the dense STP the following variant of Rayward-Smith heuristic (or the greedy algorithm [10]).

Algorithm DSTP

$(0)$ $SP \leftarrow \emptyset$;
$
\mathcal{C} \leftarrow \{\{s\} : s \in S\}$
$
(1)$ while $|\mathcal{C}| > \frac{1}{\epsilon}$ do

find $v \in V \setminus S$ with the maximum size of

$D(v) = \{C \in \mathcal{C} : C$ contains a neighbor of $v\}$

$SP \leftarrow SP \cup v$;

$
\mathcal{C} \leftarrow \mathcal{C} \setminus D(v) \cup \{\cup_{C \in D(v)} C\}$;

(2) find an optimal Steiner tree $T$ for a terminal set $S \cup SP$.

Let $\mathcal{C}$ consist of sets $C_1, ..., C_r$ after Step (1) of Algorithm DSTP. Let add edges between all terminals of the same set $C_i, i = 1, ..., r$. The length of the optimal Steiner tree in the graph $G'$ obtained cannot be longer than in $G$. There is an optimal Steiner tree $OPT'$ in $G'$ containing spanning trees $M_i$ for each set $C_i, i = 1, ..., r$. If we contract any such tree $M_i$ to a vertex, then $OPT'$ appears to be an optimal Steiner tree $M_0$ spanning vertices corresponding to $C_i$. Thus, the edge set of $OPT'$ is a union of edges of $M_i, i = 0, 1, ..., r$.

Algorithm DSTP constructs some Steiner trees $M_i'$ in $G$ for terminals of $C_i$ (step (1)) and then finds the shortest tree $M_0'$ spanning $M_i', i = 1, ..., r$ (step (2)). $M_0'$ cannot be longer that $M_0$, since $M_0$ also spans $M_i'$. Remark 1 implies that an exhaustive search in Step (2) can be executed in time $O(n^{1/\epsilon})$.

An approximation ratio of Algorithm DSTP is at most

$$\frac{\sum_{i=0}^{r} d(M_i')} {\sum_{i=0}^{r} d(M_i)} \leq \frac{\sum_{i=1}^{r} d(M_i)} {\sum_{i=1}^{r} d(M_i)} = \frac{k - r + |SP|}{k - r} \leq 1 + \frac{|SP|}{k - \frac{1}{\epsilon}}.$$ (1)
The size of $SP$ equals to the number of iterations in Step (1). Each iteration of (1) decreases the size of $\mathcal{C}$ by at least $\epsilon |\mathcal{C}| - 1$. Thus, after $i$-th iteration $|\mathcal{C}| \leq (k - \frac{1}{\epsilon})((1 - \epsilon)^i + \frac{1}{\epsilon})$. The procedure (1) interrupts when $|\mathcal{C}| < \frac{1}{\epsilon} + 1$, so

$$|SP| \leq \log_{1/(1-\epsilon)}(k - \frac{1}{\epsilon}).$$

Thus, (1) implies the following

**Lemma 3** An approximation ratio of Algorithm DSTD is at most

$$1 + \frac{\log_{1/(1-\epsilon)}(k - \frac{1}{\epsilon})}{k - \frac{1}{\epsilon}}.$$  

Given an arbitrary approximation ratio $1 + \gamma$, $\gamma > 0$, our strategy is to solve exactly in polynomial time (for fixed $\epsilon$ and $\gamma$) instances of DSTD with small number of terminals, i.e. when $k$ satisfies the following inequality

$$\frac{\log_{1/(1-\epsilon)}(k - \frac{1}{\epsilon})}{k - \frac{1}{\epsilon}} \leq \gamma.$$  

If the number of terminals is sufficiently big, then we apply Algorithm DSTD. Thus we obtain the following

**Theorem 3** There is a polynomial-time approximation scheme for the $\epsilon$-dense STP.

It is not difficult to see that there is a polynomial time reduction of the $\epsilon$-dense SCP to the $\epsilon$-dense STP and vice versa, thus, the problem of polynomial time solvability of $\epsilon$-dense STP is equivalent to Problem 1.

Similarly to SCP, we define $\delta$-superdense STP to be the case of STP where any terminal has at least $n - o(n^\delta)$ neighbors outside $S$.

**Corollary 1** The $\delta$-superdense STP can be solved exactly in polynomial time.

### 3 Dense Vertex Cover Problem

**Vertex Cover Problem (VCP).** Given a graph $G = (V, E)$, find a minimum size vertex set $OPT \subseteq V$ such that at least one end of any edge belongs to $OPT$.

The following algorithm is suggested for VCP in $\epsilon$-dense graphs, i.e., in graphs where any vertex has at least $\epsilon n$ neighbors for some $\epsilon > 0$ ($|V| = n$). Let $O(v)$ denote the set of neighbors of a vertex $v$, $G(V')$ denote a subgraph induced by a vertex set $V' \subseteq V$ and $2VC$ denote the well-known 2-approximation algorithm for VCP.
Algorithm DVC

for all $v \in V$
do $V' \leftarrow V \setminus (O(v) \cup \{v\})$
find a vertex cover $VC(v)$ for $G(V')$ using 2VC;
$VC(v) \leftarrow O(v) \cup VC(v)$;
$APPR \leftarrow \arg \min_{v \in V} |VC(v)|$.

Let $v \notin OPT$. Then $O(v) \subseteq OPT$ since all edges incident to $v$ should be covered by $OPT$. Moreover, $O(v)$ covers all edges between $O(v)$ and the corresponding $V'$. So the rest of vertices of $OPT$ cover the edges of $G(V')$.

Let $OPT'' = OPT - O(v)$. The output vertex cover of 2VC applied to $V'$ has a size at most $\min\{2|OPT''|, |V'|\}$. So the approximation ratio can be bounded as follows.

$$\frac{|APPR|}{|OPT|} \leq \frac{|O(v)| + \min\{2|OPT''|, |V'|\}}{|O(v)| + |OPT'|} \leq \min\left\{\frac{|O(v)| + 2|OPT''|}{|O(v)| + |OPT'|}, \frac{n}{|O(v)| + |OPT'|}\right\}$$

If $2|OPT'| \leq (1 - \epsilon)n$, then

$$\frac{|APPR|}{|OPT|} \leq \frac{\epsilon n + 2|OPT'|}{\epsilon n + |OPT'|} = 2 - \frac{1}{1 + \frac{|OPT'|}{\epsilon n}}$$

Thus, the more $|OPT'|$ corresponds to the more bound for the approximation ratio. Therefore,

$$\frac{|APPR|}{|OPT|} \leq 2 - \frac{1}{1 + \frac{0.5(1 - \epsilon)n}{\epsilon n}} = \frac{2}{1 + \epsilon}.$$ 

If $2|OPT'| \geq (1 - \epsilon)n$, then we obtain the same bound for the approximation ratio as follows

$$\frac{|APPR|}{|OPT|} \leq \frac{\epsilon n}{\epsilon n + 0.5(1 - \epsilon)n} = \frac{2}{1 + \epsilon}.$$ 

**Theorem 4** The algorithm DVC has an approximation ratio at most $\frac{2}{1 + \epsilon}$ for $\epsilon$-dense graphs.

**Theorem 5** The $\epsilon$-dense Vertex Cover Problem is MAX SNP-hard.

**Proof.** (Sketch.) Starting with an instance of the Vertex Cover Problem in a graph $G$ with $n$ vertices we encryp it joining all vertices of a clique of size $\frac{\epsilon}{1 - \epsilon}n$ with all vertices of $G$. The resulting graph is $\epsilon$-dense and, therefore, if we have an $\alpha$-approximation for DVC, then the reduction above gives $\alpha(1 + \epsilon)$-approximation algorithm for the general problem which is MAX SNP-hard. ◇
Further densification (as for SCP and STP) leads to decrease in the complexity of approximation.

We say that an instance of VCP is \( \delta \)-superdense if the degree of any vertex is at least \( n - o(n^\delta) \). Theorem 4 implies

**Corollary 2** The \( \delta \)-superdense VCP has a polynomial-time approximation scheme.

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**References**


