

UNIVERSITY OF CALIFORNIA
RIVERSIDE

Approximation Algorithms for the Fault-Tolerant Facility Placement Problem

A Dissertation submitted in partial satisfaction
of the requirements for the degree of

Doctor of Philosophy

in

Computer Science

by

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June 2013

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Acknowledgments

I would thank my advisor, Marek Chrobak, for bringing me into the PhD program of University of California Riverside, and for his guidance and patience on my study and research in the past five years. I am also grateful for the committee, Tao Jiang, Stefano Lonardi, and Neal Young for helpful discussions and comments on my research and this dissertation.

The supportive environment of the algorithm lab and computer science department has made PhD study here a pleasant experience and I am grateful for Claire Huang, Wei Li and the algorithm lab for helpful discussions and stimulation of ideas.

To my parents, who always have faith on my endeavor.

ABSTRACT OF THE DISSERTATION

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Doctor of Philosophy, Graduate Program in Computer Science
University of California, Riverside, June 2013
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The dissertation concerns the fault-tolerant facility placement problem (FTFP), a variant of the well-known uncapacitated facility location problem (UFL). In the FTFP problem, we have a set of sites where we can open facilities and a set of clients each with an integral demand. To satisfy their demands, clients need to be connected to open facilities in sites. The goal is to satisfy all clients' demand while minimizing the total cost, that is the cost of opening facilities and the cost of connecting clients to facilities. The problem is shown to be NP-hard and hence we study the approximation algorithms and their performance guarantee. Approximation algorithms are algorithms that run in polynomial time with provable performance when compared to optimal solutions.

In this thesis we give two techniques that lead to several LP-rounding algorithms with progressively improved approximation ratio. The best ratio we have is 1.575. This ratio matches the best LP-based approximation ratio for the more restricted problem, namely UFL. We have also studied the applicability of primal-dual approaches to FTFP. In particular, we show that a natural greedy algorithm analyzed using dual-fitting technique gives

an upper bound of $O(\log n)$ for approximation ratio. On the negative side, under a natural assumption, we give an example showing the dual-fitting analysis cannot give a ratio better than $\Omega(\log n / \log \log n)$.

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Chapter 1

Introduction

1.1 The Problem and the Background

The facility location problems (FL) are about selecting a set of candidate places to build (or open) facilities and connecting clients to the facilities to satisfy their demands. They model real world scenarios like setting up warehouses to deliver commodities to retailers, building power suppliers to serve the needs of a district of residents, placing content servers in a network to send files to clients. This simple model has been studied extensively since 1960s (see books by Mirchandani and Francis [22]) due to its practical significance and nice properties that allow a multitude of approaches to be applicable. The problems we study in this thesis fall under the category of the so-called *discrete* facility location problems. In the discrete facility location problems, we are given a set of facilities that could be opened, and a set of clients each with some demand, and the distance (or connection cost) between a client and a facility. A client needs to be serviced by getting connected to

as many open facilities as its demand. We would like to service all clients while minimizing the total cost of opening facilities and connecting clients to facilities.

The facility location problems have taken a central place in both operations research and theoretical computer science since the 60's, and a number of approaches have been proposed, including heuristic solutions, branch and bound, probabilistic methods, and more recently, approximation algorithms. Interestingly, some of the heuristic solutions, although are not concerned with the performance on the hardest instance, have later been shown to offer a provably worst-case performance guarantee.

As many optimization problems arising in practical applications, it comes without much surprise that the facility location problems is NP-hard, and thus precludes a promise for polynomial time algorithms that solve the problems exactly. Nonetheless there are polynomial time algorithms that deliver a solution with cost only a small percentage off from the cost of optimal solutions. These type of algorithms, known as approximation algorithms and their performance analysis, is the subject of this thesis.

There are over a dozen different variation in the problem formulation, and the uncapacitated facility location problem (UFL) concerns the simplest model where each candidate facility point has an opening cost and the connection cost between a client and a facility is the distance between the two. Each client needs to be connected to one open facility. A solution consists of a set of facilities to open and a specification of connections for each client to an open facility. The cost of the solution is simply the facility cost, defined as the sum of opening cost for the facilities chosen to open, and the connection cost, defined as the cost to connect every client to some open facility. The UFL problem asks for a solution

that satisfies all clients and with minimum total cost. The UFL problem with general distances has an algorithm with approximation ratio $O(\log n)$ where n is the number of clients, due to Hochbaum [16]. A matching lower bound of $O(\log n)$ is immediate, as the UFL problem contains the well-known Set-Cover problem as a special case. In the Set-Cover problem, we are given a universe $\mathcal{U} = \{e_1, \dots, e_n\}$ and a collection of sets $\mathcal{S} = \{S_1, \dots, S_m\}$ such that every $S_i, i = 1, \dots, m$ is a subset of \mathcal{U} . The problem asks for the smallest number of sets from \mathcal{S} whose union is \mathcal{U} . More on the relation of the two problems, UFL and Set-Cover in Section 1.4.

Renewed interest in the UFL problem has been made possible after Shmoys, Tardos and Aardal [23] showed that, when distances form a metric, there is an algorithm with $O(1)$ approximation ratio. Improved algorithms with more sophisticated ideas have been proposed. The past two decades has observed a sequence of improved approximation ratio, from the original 4-approximation [23], to the 1.488-approximation by Li [20], the best known approximation ratio to date.

The problem studied in this thesis is a generalization of the UFL problem, in that each client may specify a demand and the client then needs to be connected to several different facilities with the number of connections equal to its demand. A solution consists of a specification of the number of facilities to open in each site, and the number of connections between clients and sites. Open multiple facilities in the same site incurs a cost of the opening cost for this site multiplied by the number of facilities opened. The connection cost between a site and a client is the number of connections times the distance between the two, with the constraint that the number of connections cannot exceed the number of

facilities opened in that site. The FTFP problem asks for a solution with minimum total cost, that is, the sum of facility cost and connection cost.

1.2 Notation and Definition

We denote the set of sites as \mathbb{F} and the set of clients as \mathbb{C} . Each client $j \in \mathbb{C}$ has a demand r_j , meaning the client j needs to be connected to r_j different facilities. The distance between a site i and a client j is denoted as d_{ij} . To open one facility at site i incurs a cost of f_i . To make one connection from client j to a facility at site i incurs a cost of d_{ij} . The problem asks for a vector of (\mathbf{x}, \mathbf{y}) such that $x_{ij} \in \{0, 1, 2, \dots\}$ denotes the number of connections between site i and client j , and $y_i \in \{0, 1, 2, \dots\}$ denotes the number of facilities opened at site i . We then seek a solution such that $y_i \geq x_{ij}$ for every $i \in \mathbb{F}, j \in \mathbb{C}$ and $\sum_{i \in \mathbb{F}} x_{ij} \geq r_j$ for all clients $j \in \mathbb{C}$, and we are to minimize the total cost of the solution, that is $\sum_{i \in \mathbb{F}} f_i y_i + \sum_{i \in \mathbb{F}, j \in \mathbb{C}} d_{ij} x_{ij}$. We call the first term $\sum_{i \in \mathbb{F}} f_i y_i$ the facility cost of the solution, and the second term $\sum_{i \in \mathbb{F}, j \in \mathbb{C}} d_{ij} x_{ij}$ the connection cost of the solution (\mathbf{x}, \mathbf{y}) .

1.3 The Notion of P vs NP, Approximation

All problems studied in this thesis are optimization problems. An optimization problem is defined by a set of parameters and constraints. Specifying the values of the parameters defines an instance of the problem. An optimization problem is either a minimization problem or a maximization problem. A feasible solution to an optimization problem is a solution that satisfies all the constraints. There is also a cost function evaluates

every feasible solution to a numerical cost. For a minimization problem, an optimal solution is a feasible solution with the cost that is no more than the cost of any feasible solution.

An algorithm is said to solve an optimization problem exactly if for every instance, the algorithm always computes an optimal solution, or reports that no feasible solution exists. For problems that are NP-hard, such an exact algorithm that runs in polynomial time is not possible unless $P = NP$. Therefore we focus our attention on polynomial time algorithms that computes a feasible solution with a cost can be proved to be within some factor away from the cost of an optimal solution.

1.4 Hardness Results on Approximating UFL

Since FTFP contains UFL as a special case, any hardness result obtained on UFL remains applicable to FTFP. In the following, we review some well-known hardness results on UFL, with the implication that the same claims hold for FTFP as well.

The UFL problem is easily seen to be NP-hard, as they contain the Set Cover problem as a special case. The Set Cover problem is that, given a universe $\mathcal{U} = \{e_1, \dots, e_n\}$ and a collection of subsets $\mathcal{S} = \{S_1, \dots, S_m\}$ such that $S_l \leq U$ for $l = 1, \dots, m$, find a minimum number of sets in \mathcal{S} to cover all elements in \mathcal{U} . It is well-known that the Set-Cover problem is NP-hard.

Proposition 1 *The general UFL problem is NP-hard.*

Proof. Reduction from the Set Cover problem. In the Set Cover problem, we have a universe of elements, that is $\mathcal{U} = \{e_1, \dots, e_n\}$, and a collection of sets $\mathcal{S} = \{S_1, \dots, S_m\}$

such that $S_i \subseteq \mathcal{U}$ for $i = 1, \dots, m$. We construct a general UFL instance like this: for each $e_j, j = 1, \dots, n$ we have a client j , and for each set $S_i, i = 1, \dots, m$ we have a facility i . The facility cost $f_i = 1$ for every facility $i = 1, \dots, m$ ¹ and the distance $d_{ij} = 1$ if $e_j \in S_i$ and $d_{ij} = \infty$ if $e_j \notin S_i$. Clearly an optimal solution for the UFL instance can only use edges with $d_{ij} = 1$. It is easy to see that given any optimal solution of the Set Cover instance, we can construct an optimal solution for the UFL instance, by simply taking the facilities whose corresponding sets are chosen in the set cover. On the other hand, given an optimal solution to the UFL instance, we can only have $d_{ij} = 1$ connections, that implies for every client j , the corresponding element e_j is covered by some set S_i , whose corresponding facility i is chosen in the UFL solution. Let I be the set of facilities chosen in the UFL solution, it is easily seen that the corresponding sets in the Set Cover instance form a set cover. ■

Proposition 2 *The metric UFL problem is NP-hard.*

Proof. The reduction is also from the Set Cover problem. Unlike the general UFL problem, we can no longer have edges with length 1 and ∞ now, as the distances are constrained by the triangle inequality. We still have sets S_i in the Set Cover instance correspond to the facilities i in the metric UFL instance, and elements e_j in the Set Cover instance correspond to the clients j in the metric UFL instance. Our facility cost $f_i = \epsilon$ for some small number $\epsilon > 0$, and every facility i has the same facility cost. Our distance d_{ij} is now 1 if $e_j \in S_i$ and 3 if $e_j \notin S_i$ in the Set Cover instance.

Given the construction, it is clear that any optimal solution for the metric UFL instance cannot use an edge of distance 3, as there exists a solution that beats such a solution

¹Actually any value of $f_i > 0$ will work, for example, we can set $f_i = 100$ for every facility i .

with lower cost, namely a solution that opens all facilities with total cost $m\epsilon + n \cdot 1 = m\epsilon + n$, as we can choose $\epsilon = 1/m^2$. It follows that any optimal solution for the metric UFL instance must have all clients connected at distance of 1. Such a solution would have a facility set corresponds to a set cover for the corresponding Set Cover instance, as for every client j , the corresponding element e_j is covered by some set S_i corresponding to some facility i . ■

Now we show the MaxSNP-hardness of the metric UFL problem. This implies that there exists some constant c such that the metric UFL problem cannot be approximated to be better than c -approximation. As a consequence, the metric UFL problem cannot have polynomial approximation scheme (PTAS), that is, a family of algorithms that computes a solution with cost no more than $(1 + \epsilon)$ from the optimal, for any given constant $\epsilon > 0$, and the running time is polynomial in the input size with ϵ being constant.

Proposition 3 *The metric UFL problem is MaxSNP-hard [12].*

Proof. The proof is by a reduction from the B-Vertex Cover problem. In the B-Vertex Cover problem, we are given a graph $G = (V, E)$, and a constant B , such that every vertex $u \in V$ has degree no more than B . And the problem asks for a vertex cover with minimum size. That is, we are to find a minimum set $V' \subseteq V$ such that every edge $e \in E$ has at least one endpoint in V' .

The idea is to show that, for any given constant $\epsilon < 1$, given an algorithm for the metric UFL problem with approximation ratio $1 + \epsilon$ for any constant $\epsilon < 1$, we are able to find an algorithm for the B-Vertex Cover problem with approximation ratio $1 + \epsilon'$ such that ϵ' is a constant depending on ϵ and possibly B , and ϵ' approaches 0 as ϵ approaches 0. In the following we shall see that we can set $\epsilon' = (1 + B)\epsilon$ for our purpose.

Given an instance of B-Vertex Cover, we construct an instance of metric UFL. For every vertex $u \in V$ we have a facility i and for every edge $e \in E$ we have a city j . $c_{ij} = 1$ if the corresponding edge e of client j is incident on vertex u , which corresponds to the facility i , $d_{ij} = 3$ otherwise. We postpone defining the facility cost f_i but remark that all facilities i have the same facility opening cost, that is f_i equal for all facility i , so in the following we use f to denote the facility opening cost for a single facility.

Given an instance of B-Vertex Cover, let k be the size of an optimal vertex cover. The reason we need the knowledge of k is that we shall set our f as a function of k . On the other hand, we really do not know k , as it is NP-hard to compute k , the optimal solution value for B-Vertex Cover. However, we can still proceed assuming knowing k , because we can perform the following steps for every possible value of $k = 1, 2, \dots, n$ and if our proof goes through for every choice of k , our claim must hold on the (unknown) value of k as well. From now on we assume the knowledge of k . The same trick is used in the proof of Theorem 4 as well.

We are to show that we can use an $(1 + \epsilon)$ -approximation algorithm \mathcal{A}_{UFL} for metric UFL to construct an $(1 + \epsilon')$ -approximation algorithm \mathcal{A}_{BVC} for B-Vertex Cover. First we run algorithm \mathcal{A}_{UFL} on the UFL instance, and let there be βk facilities open and γn cities connect with $d_{ij} = 1$ and the rest with $d_{ij} = 3$. Then the cost of \mathcal{A}_{UFL} is

$$\text{ALG}_{\text{UFL}} = \beta k f + \gamma n + 3(1 - \gamma)n$$

Since \mathcal{A}_{UFL} is a $(1 + \epsilon)$ -approximation algorithm, we have that $\text{ALG}_{\text{UFL}} \leq (1 + \epsilon)\text{OPT}_{\text{UFL}}$. To get a handle on OPT_{UFL} , we use a feasible solution to this UFL instance. One possible choice is to use the solution with k -facilities and all clients are connected at

distance 1. Notice that we only need to know such a solution exists. This solution has cost $kf + n$. Therefore, we have

$$\beta kf + \gamma n + 3(1 - \gamma)n \leq (1 + \epsilon)(kf + n)$$

Cancelling n from both sides, we get

$$\beta kf + 2(1 - \gamma)n \leq (1 + \epsilon)kf + \epsilon n \tag{1.1}$$

Now we look for a solution to the B-Vertex Cover instance. The solution to UFL shows that we can use βk facilities to connect to γn clients with distance 1, for the others at distance 3, we need at most one facility each to connect them at distance 1, that is, we need at most $(1 - \gamma)n$ vertices to cover the remaining edges. So we have a vertex cover with size $\beta k + (1 - \gamma)n$, we hope to show that

$$\beta k + (1 - \gamma)n \leq (1 + \epsilon')k, \tag{1.2}$$

where ϵ' is a constant depend on ϵ and possibly B , with the property that ϵ' approaches 0 as ϵ approaches 0.

Compare Eqn. (1.1) and Eqn. (1.2), we need to have some way to clean up the variables to get simpler inequalities so that we can use (1.1) to deduce (1.2). Since we have the flexibility to choose f , we use that to simplify (1.1) by setting f to be such that $n/(kf) = B$, then (1.1) becomes

$$\beta + 2(1 - \gamma)B \leq (1 + \epsilon) + \epsilon B. \tag{1.3}$$

This looks very similar to the left hand side of (1.2) now, if we can get rid of n in that left hand side. We now use the fact that every vertex in the B-Vertex Cover instance has degree

at most B , so the k vertices in an optimal solution to this B-Vertex Cover instance have sum of degree at most kB , which is no less than n , the number of edges covered, since the k vertices form a vertex cover. So we have $kB \geq n$, thus, to obtain (1.2), it suffices to show

$$\beta k + (1 - \gamma)kB \leq (1 + \epsilon')k. \quad (1.4)$$

Dividing k from both sides, our goal now is to show

$$\beta + (1 - \gamma)B \leq (1 + \epsilon'). \quad (1.5)$$

Recall that in (1.3) we have $\beta + 2(1 - \gamma)B \leq (1 + \epsilon) + \epsilon B$, if we set $\epsilon' = \epsilon(1 + B)$, we shall have the following

$$\beta + (1 - \gamma)B \leq \beta + 2(1 - \gamma)B \leq (1 + \epsilon) + \epsilon B = 1 + \epsilon(1 + B) = 1 + \epsilon',$$

where the first inequality is due to $\gamma \geq 1$, and the second is from (1.3). It is easy to see that ϵ' approaches 0 as ϵ approaches 0 and we have the desired inequality (1.5). We have thus found a $(1 + \epsilon')$ approximation algorithm for the B-Vertex Cover problem, given a $(1 + \epsilon)$ -approximation algorithm for the metric UFL problem, with the property that as ϵ approaches 0, ϵ' approaches 0 as well.

Since B-Vertex Cover has MaxSNP-hard, we conclude metric UFL is MaxSNP-hard as well. ■

After we have shown the metric UFL problem is MaxSNP-hard, we focus on the metric version from now on. In the follow, we mention UFL and FTFP without explicitly specifying that their distances form a metric, that is, d_{ij} 's are symmetric and satisfy the triangle inequality.

We briefly mention the last piece of hardness result, which is the Guha-Khuller theorem.

Theorem 4 *UFL cannot be approximated to less than 1.463 unless $\text{NP} \subseteq \text{DTIME}(n^{O(\log \log n)})$.*

Proof. The proof is by contradiction. In other words, we show that if metric UFL can be solved by a polynomial time algorithm with approximation ratio α that is less than 1.463, then we have a polynomial time algorithm with approximation ratio $1/\rho \ln n$ for some constant $\rho > 1$ for Set Cover, where n is the number of elements in the universe in the Set Cover instance. Using a result by Feige [11], the existence of an $1/\rho \ln n$ approximation algorithm for some constant $\rho > 1$ implies $\text{NP} \subseteq \text{DTIME}(n^{O(\log \log n)})$.

Given a Set Cover instance with a universe $\mathcal{U} = \{e_j : j = 1, \dots, n\}$ of elements and a family of sets $\mathcal{S} = \{S_i, i = 1, \dots, m\}$ with every set $S_i \in \mathcal{S}$ being a subset of \mathcal{U} . Then the proof proceeds in iterations. In each iteration we construct a metric UFL instance with the set of facilities \mathbb{F} corresponding to the set \mathcal{S} , and we have one client j for each uncovered element e_j in \mathcal{U} . The distances are defined as 1 if $e_j \in S_i$, and 3 otherwise. We then run the supposed ρ -approximation algorithm for the constructed UFL instance. Our construction ensures that, if the UFL solution does not cover a large portion of the clients, then the ratio α between the UFL solution and an optimal integral solution for the UFL instance must be at least 1.463. The other case is that in all iterations we have an UFL solution that covers a large portion of clients, and this gives us a solution for the given Set Cover instance with no more than $1/\rho \ln n$ times of an optimal solution for some $\rho > 1$, which then implies $\text{NP} \subseteq \text{DTIME}(n^{O(\log \log n)})$.

Let k be the number of sets in an optimal solution for the Set Cover instance.

Note that we can run the following for every $k = 1, \dots, n$ so we can assume we know k . We now give the construction of the UFL instance. In iteration t , suppose we begin with n_t elements uncovered. Then we have a UFL instance with $|\mathbb{F}| = m, |\mathbb{C}| = n_t, d_{ij} = 1$ or 3 and f_i to be specified but all $i \in \mathbb{F}$ have equal f_i , although f_i is set to a different value in each iteration. Now suppose the UFL algorithm chooses $\beta_t k$ facilities and covers $\gamma_t n_t$ clients, the rest $(1 - \gamma_t)n_t$ clients are then served at a distance of 3 . The cost of this UFL solution is

$$\text{ALG}_{\text{UFL}} = \beta_t k f_i + \gamma_t n_t \cdot 1 + (1 - \gamma_t)n_t \cdot 3.$$

We set $f_i = c n_t / k$ for some constant c . Notice that the same constant c is used for all iterations and we shall use c to optimize our analysis on the lower bound of approximation ratio of any UFL algorithms. The cost of the UFL solution now becomes

$$\beta_t c n_t + \gamma_t n_t + 3(1 - \gamma_t)n_t = n_t(c\beta_t + 3 - 2\gamma_t).$$

On the other hand, we know that there exists a solution with k facilities that covers all clients at distance 1 , which corresponds to an optimal Set Cover solution. So an optimal UFL solution has cost no more than

$$\text{OPT}_{\text{UFL}} \leq k f_i + n_t = c n_t + n_t = (1 + c)n_t.$$

Since we are running an α -approximation algorithm for the UFL instance, we have

$$\alpha \geq \frac{\text{ALG}_{\text{UFL}}}{\text{OPT}_{\text{UFL}}} \geq \frac{n_t(c\beta_t + 3 - 2\gamma_t)}{(1 + c)n_t} = \frac{\beta_t + 3 - 2\gamma_t}{1 + c}.$$

Let $\rho > 1$ be some fixed constant. Now we have two cases:

Case 1: There exists some iteration t such that $\gamma_t < 1 - e^{-\rho\beta_t}$. Then we have

$$\alpha > \frac{c\beta_t + 3 - 2(1 - e^{-\rho\beta_t})}{1 + c} = \frac{c\beta_t + 1 + 2e^{-\rho\beta_t}}{1 + c}.$$

Because we have no control over β_t , we set β_t so that the right hand side is minimized, and this would surely be a lower bound on α . Fix c and ρ , and view the right hand side as a function of a single variable of β_t . We then choose β_t to minimize the right hand side. By setting the derivative with respect to β_t to zero, we have $\beta_t = (1/\rho) \ln(2\rho/c)$. Therefore

$$\alpha > \frac{c/\rho \ln(2\rho/c) + 1 + c/\rho}{1 + c}.$$

Now we deal with ρ . For any fixed $c > 0$, we observe that the right hand side is a decreasing function of $\rho > 1$. So the lower bound on α is tightest when ρ is close to 1. By taking ρ arbitrarily close to 1 (although ρ remains a constant strictly greater than 1), we have $\alpha > \frac{c \ln(2/c) + 1 + c}{1 + c}$. Since we have the choice of c , we choose c to maximize the right hand side of the above, by taking the derivative of the right hand side with respect to c and set it to zero. This will give us as good a lower bound on α as possible. We thus have c being the solution of the equation $\ln(2/c) = 1 + c$ and solve for $c = 0.463$. Substitute the value c back, we have $\alpha > 1.463$ in this case.

Case 2: The other case is that we have $\gamma_t \geq 1 - e^{-\rho\beta_t}$ for every iteration t . Suppose we have l iterations, we then have

$$n(1 - \gamma_1)(1 - \gamma_2) \dots (1 - \gamma_l) = 1.$$

Applying the assumed inequality $\gamma_t \geq 1 - e^{-\rho\beta_t}$, we have

$$n \prod_{t=1}^l e^{-\rho\beta_t} \geq 1,$$

which is

$$e^{-\sum_{t=1}^l \rho\beta_t} \geq 1/n,$$

which is

$$\sum_{t=1}^l \rho \beta_t \leq \ln n.$$

Notice that we have a Set Cover solution with $(\sum_{t=1}^l \beta_t)k$ sets, so this Set Cover solution has cost no more than $\sum_{t=1}^l \beta_t$ times k , where k is the number of sets in an optimal Set Cover solution. So we have a Set Cover algorithm with approximation ratio no more than $\sum_{t=1}^l \beta_t k / k \leq 1/\rho \ln n$ for some $\rho > 1$. Using Feige's result we have that metric UFL cannot have approximation algorithm with ratio less than 1.463 unless $\text{NP} \subseteq \text{DTIME}(n^{O(\log \log n)})$.

■

Using an observation by Sviridenko, the underlying assumption can be strengthened to $\text{P} \neq \text{NP}$. That is, metric UFL cannot have a polynomial time algorithm with ratio less than 1.463 unless $\text{P} = \text{NP}$. With that, we conclude the discussion on the hardness results for the Uncapaciated Facility Location problem (UFL), with the implication that all these hardness results apply to our problem, the Fault-tolerant Facility Placement problem (FTFP) as well.

Chapter 2

Related Work and Known Results

In designing algorithms for the facility location problems (FL), which encompass all three problems addressed in the following, the Uncapacitated Facility Location problem (UFL), the Fault-tolerant Facility Location problem (FTFL), and the Fault-tolerant Facility Placement problem (FTFP), we have two competing goals: On one hand we want to open as few facilities as possible so that our facility cost is small; on the other hand we need as many facilities as possible so that every client can connect to a nearby facility.

Some simple observations about the UFL problem, and apply to other FL problems with minor adjustments:

Observation 5 *If all facility opening cost are zero, then an optimal solution is to open all facilities and connect each client to the nearest facility.*

Observation 6 *If all distances are zero, that is, all facilities and clients are colocated at the same point, then an optimal solution is to open the cheapest facility and connect all clients to that facility.*

Observation 7 *For a fixed set of open facilities, there is a polynomial algorithm to find the optimal assignment from clients to those facilities.*

Proof. For UFL, it is a simple matter of finding the nearest open facility for each client and a brute force search only takes time $O(|\mathbb{F}| \cdot |\mathbb{C}|)$. Even for more complicated variants where the facilities each has a capacity, the optimal assignment can be found by using the mincost flow algorithm. ■

Next we review the known algorithms for UFL and FTFL, since these two are well studied problems in literature. In particular, the LP-rounding algorithms for UFL inspired our approach to the FTFL problem.

2.1 Related Work on UFL

The Uncapacitated Facility Location problem (UFL) is the simplest variant of the Facility Location problems (FL), and has received the most attention in research community. One thing that is surprising about UFL is that almost all known approximation algorithm design techniques can be applied to UFL with a good approximation ratio. The first $O(1)$ -approximation algorithm was obtained by Shmoys, Tardos and Aardal [23], with a ratio of 3.16, using LP-rounding. Following that, more sophisticated rounding algorithms have been proposed with improved approximation ratios. Significant improvement comes from randomized rounding and Chudak and Shmoys [8], and Sviridenko [24] achieved a ratio of 1.736 and 1.582 respectively. The LP-rounding algorithms require solving the LP. More efficient algorithms that are combinatorial have been proposed as well. Jain and Vazirani [18] introduced a primal-dual algorithm which grows a feasible dual solution and

updates a primal solution accordingly until the primal solution is feasible. The approximation ratio is obtained via a relaxed version of the complementary slackness conditions. Two greedy algorithms that repeatedly picking the star with minimum average cost, where a star is a facility and a subset of clients, has been analyzed by Jain, Markakis, Mahdian, Saberi and Vazirani [17] using the dual-fitting technique. They showed that the cost of the integral solution is equal to the sum of dual variable values produced by the greedy algorithm, and the dual variable values, after shrinking by a common factor, is feasible to the dual program. That common factor is then the desired approximation ratio. They showed their algorithms have approximation ratio 1.861 and 1.61 respectively. Another approach is local search, in which we start with a feasible integral solution and make local moves to improve the solution, and stop at some local optimum. Arya *et al.* [1] showed that local search gives a ratio of 3 for the UFL problem.

Given that the total cost of a solution to the UFL problem consists of two parts, the facility cost and the connection cost, a notion of bifactor approximation was introduced in the literature. An algorithm with facility cost F_{ALG} and connection cost C_{ALG} , is said to be (γ_f, γ_c) approximation if, for every feasible solution SOL, with facility cost F_{SOL} and connection cost C_{SOL} , we have

$$F_{\text{ALG}} + C_{\text{ALG}} \leq \gamma_f F_{\text{SOL}} + \gamma_c C_{\text{SOL}}.$$

In particular, for the optimal fractional solution $(\mathbf{x}^*, \mathbf{y}^*)$ with facility cost $F^* = \sum_{i \in \mathbb{F}} f_i y_i^*$ and connection cost $C^* = \sum_{j \in \mathbb{C}} d_{ij} x_{ij}^*$, we have

$$F_{\text{ALG}} + C_{\text{ALG}} \leq \gamma_f F^* + \gamma_c C^*.$$

The notion of bifactor approximation is helpful when an algorithm has imbalanced factors γ_f and γ_c . It is easy to see such an algorithm has approximation ratio $\max\{\gamma_f, \gamma_c\}$. However, more can be said, as there are techniques like cost scaling and greedy augmentation to balance the two factors, that is, to decrease one at the expense of increasing the other, and achieve a better overall approximation ratio. The techniques are introduced by Guha, Khuller and Charikar [12, 7].

We now present a more detailed description on the LP-rounding approaches, as our results on FTFP are built on the work of LP-rounding for UFL.

Every LP-rounding algorithm for UFL starts with solving the LP to obtain an optimal fraction solution $(\mathbf{x}^*, \mathbf{y}^*)$. Then we need to round the fractional solution to an integral solution $(\hat{\mathbf{x}}, \hat{\mathbf{y}})$ without increasing the cost by much. As hinted in the above discussion, an integral solution with small cost would have each client connect to a nearby facility and few facilities open. Consider a client j , to get a handle on the connection cost, we would like j to connect to some neighboring facility $i \in N(j)$, where $N(j) = \{i \in \mathbb{F} : x_{ij}^* > 0\}$. However, it is in general not possible to have every client connect to a neighboring facility, or we would open too many facilities, and thus incur a high facility cost. An alternative is to select a subset of clients, denoted as $C' \subseteq \mathbb{C}$ and only require clients in C' have a neighboring facility open. Clients outside C' are then connected to a facility via some clients in C' . The connection cost for clients in $\mathbb{C} \setminus C'$ are then bounded using the triangle inequality.

Suppose we have chosen a subset C' of client, we then open one facility for each client j' in C' and j' would connect to that facility. Denote by $\phi(j)$ the facility that client j connects to. To bound the facility cost F_{ALG} , it suffices to require the clients j' in C' have

disjoint neighborhood $N(j')$, and $f_{\phi(j')}$ be bounded by the average facility cost of $N(j')$, that is, $f_{\phi(j')} \leq \sum_{i \in N(j')} f_i y_i^*$. Call two clients j_1 and j_2 *related* if their neighborhood overlap, that is $N(j_1) \cap N(j_2) \neq \emptyset$. It is easy to see this relation defines an equivalence class. The requirement of disjoint neighborhood on the set C' immediately implies that from each equivalence class we can select at most one client into set C' . The chosen client is then called the *representative* of that class, and we say other clients in that class are *assigned* to that representative.

Suppose we do select exactly one representative client from each equivalence class, and we open one facility, which is $\phi(j')$ for each representative client j' in C' in a way that $f_{\phi(j')}$ is no more than the average cost of facilities in $N(j')$, that is $f_{\phi(j')} \leq \sum_{i \in N(j')} f_i y_i$. Then we know that the facility cost is bounded by $\sum_{i \in \mathbb{F}} f_i y_i^* = F^*$ because the neighborhood of clients in C' are disjoint.

For connection cost, each client j' in C' connects to the only facility open in $N(j')$ and we have mentioned that there is a way to bound the distance from $\phi(j')$ to j' when $\phi(j') \in N(j')$. For clients j not in C' , we need to have an open facility to connect to, in case none of its neighbor opens. One possibility is to use $\phi(j')$, if j is assigned to j' . When that happens, $d_{\phi(j'),j}$ can be bounded by the triangle inequality as

$$d_{\phi(j),j} = d_{\phi(j'),j} \leq d_{\phi(j'),j'} + d_{j'j} \leq d_{\phi(j'),j'} + d_{i,j'} + d_{i,j} \text{ for all } i \in \mathbb{F}.$$

In the above $d_{jj'}$ denotes the shortest distance between client j and j' , that is $d_{jj'} = \min_{i \in \mathbb{F}} d_{ij'} + d_{ij}$. Now if we can show a bound on both $d_{\phi(j'),j'}$ and $d_{j'j}$ in terms of some quantity of client j , we can expect to have some bound on $d_{\phi(j'),j}$.

Since the above holds for any facility $i \in \mathbb{F}$, it would certainly hold for some

facility i' in $N(j') \cap N(j)$. Such facility i' exists because the way the relation between clients are defined. We then have a bound on all three quantities, namely $d_{\phi(j'),j'}$, $d_{i'j'}$ and $d_{i'j}$ because the corresponding facilities are all in the clients' neighborhood. Now we only need $d_{\phi(j'),j'}$ and $d_{i'j'}$ to be small in terms of some bound on d_{ij} . As we have a choice on which representative we pick in each equivalence class, we pick the representative to minimize some measure on that client. Let $g(j)$ be some measure on client j such that for any facility i in $N'(j)$, we have $d_{ij} \leq g(j)$, where $N'(j)$ is a subset of facilities that depends on the definition of $g(j)$. Then our integral solution consists of one facility from each $N'(j')$. Given $g(j)$ for all clients $j \in \mathbb{C}$, we can choose the representative j' as the one that minimizes the measure $g(j)$ in its equivalence class. One consequence is that any client j outside C' , can now have its connection cost being bounded as:

$$d_{\phi(j),j} \leq d_{\phi(j'),j'} + d_{i'j'} + d_{i'j} \leq g(j') + g(j') + g(j) \leq g(j) + g(j) + g(j) = 3g(j).$$

In other words, given the function $g(j)$ and its associated subsets of facilities $N'(j)$ for each client $j \in \mathbb{C}$, we are able to round the fractional solution $(\mathbf{x}^*, \mathbf{y}^*)$ into an integral solution such that $F_{\text{ALG}} \leq F^*$ and $d_{\phi(j),j} \leq 3g(j)$.

Next we discuss approaches to derive the function $g(j)$. Two ways are possible: One is to use the dual solution $(\boldsymbol{\alpha}^*, \boldsymbol{\beta}^*)$, and then the complementary slackness conditions tell us that $d_{ij} \leq \alpha_j^*$ for every $i \in N(j) = \{i \in \mathbb{F} : x_{ij}^* > 0\}$. This defines $g(j) = \alpha_j^*$ and $N'(j) = N(j)$. The other way is to trim the neighborhood of a client by throwing away most distant neighbors so that the farthest facility in the trimmed neighborhood is at a distance within some factor of the average distance $C_j^* = \sum_{i \in N(j)} d_{ij} x_{ij}^*$. We now open one facility in $N'(j')$ for each representative $j' \in C'$ in a way that $f_{\phi(j')}$ is no more than the

average facility cost of facilities in $N'(j')$, that is $\sum_{i \in N'(j')} f_i y_i / \sum_{i \in N'(j')} y_i$. Now we have a bound on $d_{\phi(j),j}$ in terms of C_j^* . However, our facility cost is now bounded by

$$F_{\text{ALG}} \leq \sum_{j' \in C'} \left(\sum_{i \in N'(j')} f_i y_i^* / \sum_{i \in N'(j')} y_i^* \right).$$

To estimate this cost, we assume a lower bound γ on $\sum_{i \in N'(j')} y_i^*$ for every client $j' \in C'$.

Our facility cost can then be bounded as

$$\begin{aligned} F_{\text{ALG}} &\leq \sum_{j' \in C'} \left(\sum_{i \in N'(j')} f_i y_i^* / \sum_{i \in N'(j')} y_i^* \right) \quad (\text{choice of } \phi(j')) \\ &\leq \sum_{j' \in C'} \sum_{i \in N'(j')} f_i y_i^* / \gamma \quad (\text{assumption } \sum_{i \in N'(j')} y_i^* \leq \gamma) \\ &= \frac{1}{\gamma} \sum_{j' \in C'} \sum_{i \in N'(j')} f_i y_i^* \\ &\leq \frac{1}{\gamma} \sum_{j' \in C'} \sum_{i \in N(j')} f_i y_i^* \quad (\text{replace } N'(j) \text{ with } N(j)) \\ &= \frac{1}{\gamma} \sum_{i \in \mathbb{F}} f_i y_i^* = F^*. \end{aligned}$$

We note that the factor that bounds $g(j)$ by C_j^* is related to γ and we denote that factor as $f(\gamma)$. Then we have $g(j) \leq f(\gamma)C_j^*$. It follows that we can obtain an integral solution with facility cost no more than γF^* and connection cost no more than $3 \sum_{j \in \mathbb{C}} g(j) = 3 \sum_{j \in \mathbb{C}} f(\gamma)C_j^* = 3f(\gamma)C^*$. Recall that $F^* = \sum_{i \in \mathbb{F}} f_i y_i^*$ and $C^* = \sum_{i \in \mathbb{F}, j \in \mathbb{C}} d_{ij} x_{ij}^*$ and we also define $\text{LP}^* = F^* + C^*$. Our rounded solution then has approximation ratio $\max\{\gamma, 3f(\gamma)\}$.

For concreteness we now describe a simple rounding by Chudak [9], and the original approach by Shmoys, Tardos and Aardal [23], followed by improvement by Chudak and

Shmoys [8] and Sviridenko [24], Byrka [3] and Li [20].

We start with an result by Chudak which gives an easy 4-approximation. The algorithm starts with fractional optimal solution for the primal $(\mathbf{x}^*, \mathbf{y}^*)$ and the dual $(\boldsymbol{\alpha}^*, \boldsymbol{\beta}^*)$. The algorithm then repeatedly pick an unconnected client j with minimum α_j^* , open a cheapest facility i in $N(j)$, and then connect remaining clients to i with a neighborhood overlapping $N(j)$. Given the above discussion, clearly the facility cost is no more than F^* . For connection cost, we can use $g(j) = \alpha_j^*$, and client j is connected to a facility with distance no more than $3g(j) = 3\alpha_j^*$. Summarizing we have total cost no more than $F^* + \sum_{j \in \mathbb{C}} \alpha_j^* = F^* + 3\text{LP}^* \leq 4\text{LP}^*$.

Another 4-approximation, which is the original approach by Shmoys, Tardos and Aardal uses the other way to obtain $g(j)$, by cutting the neighborhood. More precisely, for each client j , they cut at a distance $g(j)$ such that the accumulated fractional value is at least some constant γ , that is

$$\sum_{i \in N(j) : d_{ij} \leq g(j)} x_{ij}^* \geq \gamma.$$

An argument similar to Markov's inequality would then show that the trimmed neighborhood, call it $N'(j)$, has the farthest facility within a distance of $g(j)$. To bound $g(j)$, we have

$$C_j^* \stackrel{\text{def}}{=} \sum_{i \in N(j)} d_{ij} x_{ij}^* \geq \sum_{i \in N(j) \setminus N'(j)} d_{ij} x_{ij}^* \geq \sum_{i \in N(j) \setminus N'(j)} g(j) x_{ij}^* \geq g(j)(1 - \gamma).$$

As a result $g(j) \leq C_j^*/(1 - \gamma)$ and notice that $\sum_{j \in \mathbb{C}} C_j^* = C^*$. As discussed above, we have total connection cost no more than $3C^*/(1 - \gamma)$. On the facility cost, since we have limited ourselves only facilities in $N'(j)$, as opposed to $N(j)$. If we open the cheapest facility in

$N'(j)$ for each chosen client j , all we can say now is that the facility cost is not more than

$$\sum_{j \in P} \sum_{i \in N'(j)} f_i y_i^* / \sum_{i \in N'(j)} y_i^*.$$

We are blessed since we have an lower bound of γ on $\sum_{i \in N'(j)} y_i^*$ for every client j . As shown earlier, the above sum can be upper bounded by

$$\sum_{j \in P} \sum_{i \in N'(j)} f_i y_i^* / \gamma \leq \frac{1}{\gamma} \sum_{j \in P} \sum_{i \in N(j)} f_i y_i^* \leq \frac{1}{\gamma} \sum_{i \in \text{siteset}} f_i y_i^* = \frac{1}{\gamma} F^*.$$

So our approximation ratio is $\max\{3/(1 - \gamma), 1/\gamma\}$. Pick $\gamma = 1/4$ and we have a 4-approximation.

The simplest rounding described so far is clearly not the best possible and there are several improvements proposed afterwards, nonetheless, they all follow the same equivalence class partition and opens exactly one facility for each representative. To save on connection cost, one observes that a not chosen client can use a neighboring facility provided that the neighbor is open. For the interest of facility cost, it is not necessary to open the cheapest, by opening each facility in a neighborhood with probability proportional to y_i^* , we would have an expectation of facility cost bounded by F^* as well. The analysis, however, is much more involved and requires one to estimate the probabilities of a not chosen client connects to its neighbor and to the facility opened by its assigned representative client. We shall not elaborate here as our rounding algorithms for the FTFP problem in the main part will rephrase and prove similar theorems later. This concludes our review on the known approximation results on the Uncapacitated Facility Location problem (UFL).

2.2 Related Work on FTFL

The fault-tolerant facility location problem (FTFL), was first introduced by Jain and Vazirani [19] and they adapted their primal-dual algorithm for UFL to FTFL to obtain a ratio of $3 \ln R$ where $R = \max_j r_j$ is the maximum demand among all clients. The first constant approximation algorithm was given by Guha, Meyerson and Munagala [13], using LP-rounding similar to the Shmoys' *et al.*'s approach for the UFL problem. Subsequent improvement were made by Swamy and Shmoys [25] using pipage rounding with a ratio of 2.076. The current best known approximation ratio is 1.7245, by Byrka, Srinivasan and Swamy [6], using dependent rounding with a laminar clustering structure.

We note that all the known $O(1)$ -approximation algorithms for FTFL are LP-rounding algorithms and they need to solve the LP as a first step. Given the success of primal-dual based approaches for UFL, it is natural to ask whether such algorithms could be adapted to FTFL with a good ratio. To the best of the author's knowledge, it is not known whether there is a primal-dual algorithm for FTFL with a sub-logarithmic approximation ratio. We shall have more to say about this in Chapter 6.

2.3 Our Problem: FTFP

Our problem, the fault-tolerant facility placement problem (FTFP), was introduced by Xu and Shen [26]¹. The study of FTFP was partly motivated to obtain a better understanding of the implication of fault-tolerant requirement on facility location problems. The Xu and Shen's results seem to be valid only for a special case of FTFP and we later

¹In their paper they call the problem the fault-tolerant facility allocation problem, or FTFA.

adapted the Chudak's 4-approximation algorithm for UFL to FTFP, thus obtaining the first $O(1)$ -approximation algorithm for this problem. The algorithm was based on LP-rounding. Our subsequent work improved the algorithm and analysis on LP-rounding and our most recent results demonstrated that LP-rounding algorithms for FTFP can achieve an approximation ratio that matches that for UFL. For the applicability of primal-dual techniques on FTFP, we provide an argument of possible difficulty in obtaining sub-logarithmic ratio using dual-fitting approach. Again we refer the reader to Chapter 6 for more discussion on results using primal-dual approaches on FTFP.

Chapter 3

Linear Program

In all our algorithms, we use the standard linear program 3.1 and its dual 3.2, first formulated by Balinski, to develop our algorithms and estimate the optimal solution value. For readers unfamiliar with Linear Programming (LP) and Integer Programming (IP), we included a short introductory section in the appendix.

The FTFP problem has a natural Integer Programming (IP) formulation. Let y_i represent the number of facilities built at site i and let x_{ij} represent the number of connections from client j to facilities at site i . If we relax the integrality constraints, we obtain the following LP:

$$\begin{aligned}
& \text{minimize} && \text{cost}(\mathbf{x}, \mathbf{y}) = \sum_{i \in \mathbb{F}} f_i y_i + \sum_{i \in \mathbb{F}, j \in \mathbb{C}} d_{ij} x_{ij} && (3.1) \\
& \text{subject to} && y_i - x_{ij} \geq 0 && \forall i \in \mathbb{F}, j \in \mathbb{C} \\
& && \sum_{i \in \mathbb{F}} x_{ij} \geq r_j && \forall j \in \mathbb{C} \\
& && x_{ij} \geq 0, y_i \geq 0 && \forall i \in \mathbb{F}, j \in \mathbb{C}
\end{aligned}$$

The dual program is:

$$\begin{aligned}
& \text{maximize} && \sum_{j \in \mathbb{C}} r_j \alpha_j && (3.2) \\
& \text{subject to} && \sum_{j \in \mathbb{C}} \beta_{ij} \leq f_i && \forall i \in \mathbb{F} \\
& && \alpha_j - \beta_{ij} \leq d_{ij} && \forall i \in \mathbb{F}, j \in \mathbb{C} \\
& && \alpha_j \geq 0, \beta_{ij} \geq 0 && \forall i \in \mathbb{F}, j \in \mathbb{C}
\end{aligned}$$

In each of our algorithms we will fix some optimal solutions of the LPs (3.1) and (3.2) that we will denote by $(\mathbf{x}^*, \mathbf{y}^*)$ and $(\boldsymbol{\alpha}^*, \boldsymbol{\beta}^*)$, respectively.

With $(\mathbf{x}^*, \mathbf{y}^*)$ fixed, we can define the optimal facility cost as $F^* = \sum_{i \in \mathbb{F}} f_i y_i^*$ and the optimal connection cost as $C^* = \sum_{i \in \mathbb{F}, j \in \mathbb{C}} d_{ij} x_{ij}^*$. Then $\text{LP}^* = \text{cost}(\mathbf{x}^*, \mathbf{y}^*) = F^* + C^*$ is the joint optimal value of (3.1) and (3.2). We can also associate with each client j its fractional connection cost $C_j^* = \sum_{i \in \mathbb{F}} d_{ij} x_{ij}^*$. Clearly, $C^* = \sum_{j \in \mathbb{C}} C_j^*$. Throughout the paper we will use notation OPT for the optimal integral solution of (3.1). OPT is the value

we wish to approximate, but, since $\text{OPT} \geq \text{LP}^*$, we can instead use LP^* to estimate the approximation ratio of our algorithms.

Completeness and facility splitting. Define $(\mathbf{x}^*, \mathbf{y}^*)$ to be *complete* if $x_{ij}^* > 0$ implies that $x_{ij}^* = y_i^*$ for all i, j . In other words, each connection either uses a site fully or not at all. As shown by Chudak and Shmoys [8], we can modify the given instance by adding at most $|\mathbb{C}|$ sites to obtain an equivalent instance that has a complete optimal solution, where “equivalent” means that the values of F^* , C^* and LP^* , as well as OPT , are not affected. Roughly, the argument is this: We notice that, without loss of generality, for each client k there exists at most one site i such that $0 < x_{ik}^* < y_i^*$. We can then perform the following *facility splitting* operation on i : introduce a new site i' , let $y_{i'}^* = y_i^* - x_{ik}^*$, redefine y_i^* to be x_{ik}^* , and then for each client j redistribute x_{ij}^* so that i retains as much connection value as possible and i' receives the rest. Specifically, we set

$$\begin{aligned} y_{i'}^* &\leftarrow y_i^* - x_{ik}^*, \quad y_i^* \leftarrow x_{ik}^*, \quad \text{and} \\ x_{i'j}^* &\leftarrow \max(x_{ij}^* - x_{ik}^*, 0), \quad x_{ij}^* \leftarrow \min(x_{ij}^*, x_{ik}^*) \quad \text{for all } j \neq k. \end{aligned}$$

This operation eliminates the partial connection between k and i and does not create any new partial connections. Each client can split at most one site and hence we shall have at most $|\mathbb{C}|$ more sites.

By the above paragraph, without loss of generality we can assume that the optimal fractional solution $(\mathbf{x}^*, \mathbf{y}^*)$ is complete. This assumption will in fact greatly simplify some of the arguments in the paper. Additionally, we will frequently use the facility splitting operation described above in our algorithms to obtain fractional solutions with desirable

properties.

Chapter 4

Techniques

After obtaining an optimal fractional solution to LP(3.1) and (3.2), we employ two techniques to obtain approximation results on the FTFP problem. Our first technique, which we call *demand reduction*, allow us to restrict our attention to a restricted version of the FTFP problem, in which all demands r_j are not too large. This restriction then allow the application of our next technique, *adaptive partition*, so that we obtain an FTFP instance with facilities and unit demand points. For this FTFP instance, each facility can be either open or close, and each unit demand point connects to one of the open facilities. We would like to point out that we still need to cater the fault-tolerant requirement, that is, unit demands originated from the same client must connect to different facilities. We shall see that our adaptive partitioning step takes care of the fault-tolerant requirement smoothly.

4.1 Demand Reduction

This section presents a *demand reduction* trick that reduces the problem for arbitrary demands to a special case where demands are bounded by $|\mathbb{F}|$, the number of sites. (The formal statement is a little more technical – see Theorem 9.) Our algorithms in the sections that follow process individual demands of each client one by one, and thus they critically rely on the demands being bounded polynomially in terms of $|\mathbb{F}|$ and $|\mathbb{C}|$ to keep the overall running time polynomial.

The reduction is based on an optimal fractional solution $(\mathbf{x}^*, \mathbf{y}^*)$ of LP (3.1). From the optimality of this solution, we can also assume that $\sum_{i \in \mathbb{F}} x_{ij}^* = r_j$ for all $j \in \mathbb{C}$. As explained in Section 3, we can assume that $(\mathbf{x}^*, \mathbf{y}^*)$ is complete, that is $x_{ij}^* > 0$ implies $x_{ij}^* = y_i^*$ for all i, j . We split this solution into two parts, namely $(\mathbf{x}^*, \mathbf{y}^*) = (\hat{\mathbf{x}}, \hat{\mathbf{y}}) + (\dot{\mathbf{x}}, \dot{\mathbf{y}})$, where

$$\hat{y}_i \leftarrow \lfloor y_i^* \rfloor, \quad \hat{x}_{ij} \leftarrow \lfloor x_{ij}^* \rfloor \quad \text{and}$$

$$\dot{y}_i \leftarrow y_i^* - \lfloor y_i^* \rfloor, \quad \dot{x}_{ij} \leftarrow x_{ij}^* - \lfloor x_{ij}^* \rfloor$$

for all i, j . Now we construct two FTFP instances $\hat{\mathcal{I}}$ and $\dot{\mathcal{I}}$ with the same parameters as the original instance, except that the demand of each client j is $\hat{r}_j = \sum_{i \in \mathbb{F}} \hat{x}_{ij}$ in instance $\hat{\mathcal{I}}$ and $\dot{r}_j = \sum_{i \in \mathbb{F}} \dot{x}_{ij} = r_j - \hat{r}_j$ in instance $\dot{\mathcal{I}}$. It is obvious that if we have integral solutions to both $\hat{\mathcal{I}}$ and $\dot{\mathcal{I}}$ then, when added together, they form an integral solution to the original instance. Moreover, we have the following lemma.

Lemma 8 (i) $(\hat{\mathbf{x}}, \hat{\mathbf{y}})$ is a feasible integral solution to instance $\hat{\mathcal{I}}$.

(ii) $(\dot{\mathbf{x}}, \dot{\mathbf{y}})$ is a feasible fractional solution to instance $\dot{\mathcal{I}}$.

(iii) $\dot{r}_j \leq |\mathbb{F}|$ for every client j .

Proof. (i) For feasibility, we need to verify that the constraints of LP (3.1) are satisfied. Directly from the definition, we have $\hat{r}_j = \sum_{i \in \mathbb{F}} \hat{x}_{ij}$. For any i and j , by the feasibility of $(\mathbf{x}^*, \mathbf{y}^*)$ we have $\hat{x}_{ij} = \lfloor x_{ij}^* \rfloor \leq \lfloor y_i^* \rfloor = \hat{y}_i$.

(ii) From the definition, we have $\dot{r}_j = \sum_{i \in \mathbb{F}} \dot{x}_{ij}$. It remains to show that $\dot{y}_i \geq \dot{x}_{ij}$ for all i, j . If $x_{ij}^* = 0$, then $\dot{x}_{ij} = 0$ and we are done. Otherwise, by completeness, we have $x_{ij}^* = y_i^*$. Then $\dot{y}_i = y_i^* - \lfloor y_i^* \rfloor = x_{ij}^* - \lfloor x_{ij}^* \rfloor = \dot{x}_{ij}$.

(iii) From the definition of \dot{x}_{ij} we have $\dot{x}_{ij} < 1$. Then the bound follows from the definition of \dot{r}_j . ■

Notice that our construction relies on the completeness assumption; in fact, it is easy to give an example where $(\dot{\mathbf{x}}, \dot{\mathbf{y}})$ would not be feasible if we used a non-complete optimal solution $(\mathbf{x}^*, \mathbf{y}^*)$. Note also that the solutions $(\hat{\mathbf{x}}, \hat{\mathbf{y}})$ and $(\dot{\mathbf{x}}, \dot{\mathbf{y}})$ are in fact optimal for their corresponding instances, for if a better solution to $\hat{\mathcal{I}}$ or $\dot{\mathcal{I}}$ existed, it could give us a solution to \mathcal{I} with a smaller objective value.

Theorem 9 *Suppose that there is a polynomial-time algorithm \mathcal{A} that, for any instance of FTFP with maximum demand bounded by $|\mathbb{F}|$, computes an integral solution that approximates the fractional optimum of this instance within factor $\rho \geq 1$. Then there is a ρ -approximation algorithm \mathcal{A}' for FTFP.*

Proof. Given an FTFP instance with arbitrary demands, Algorithm \mathcal{A}' works as follows: it solves the LP (3.1) to obtain a fractional optimal solution $(\mathbf{x}^*, \mathbf{y}^*)$, then it constructs instances $\hat{\mathcal{I}}$ and $\dot{\mathcal{I}}$ described above, applies algorithm \mathcal{A} to $\dot{\mathcal{I}}$, and finally

combines (by adding the values) the integral solution $(\hat{\mathbf{x}}, \hat{\mathbf{y}})$ of $\hat{\mathcal{I}}$ and the integral solution of $\dot{\mathcal{I}}$ produced by \mathcal{A} . This clearly produces a feasible integral solution for the original instance \mathcal{I} . The solution produced by \mathcal{A} has cost at most $\rho \cdot \text{cost}(\dot{\mathbf{x}}, \dot{\mathbf{y}})$, because $(\dot{\mathbf{x}}, \dot{\mathbf{y}})$ is feasible for $\dot{\mathcal{I}}$. Thus the cost of \mathcal{A}' is at most

$$\text{cost}(\hat{\mathbf{x}}, \hat{\mathbf{y}}) + \rho \cdot \text{cost}(\dot{\mathbf{x}}, \dot{\mathbf{y}}) \leq \rho(\text{cost}(\hat{\mathbf{x}}, \hat{\mathbf{y}}) + \text{cost}(\dot{\mathbf{x}}, \dot{\mathbf{y}})) = \rho \cdot \text{LP}^* \leq \rho \cdot \text{OPT},$$

where the first inequality follows from $\rho \geq 1$. This completes the proof. ■

Two nice consequences are immediate, given demand reduction.

4.1.1 Reduction from FTFP to FTFL

Given demand reduction, we may assume that we are working with a restricted version of FTFP where every demand r_j is no more than $|\mathbb{F}|$. In this case we can reduce this version of FTFP into FTFL. The reduction simply creates $|\mathbb{F}|$ facilities at each site, and every such facility may be open or close later. Then we have an FTFL instance where every client have a demand r_j and every facility could be open or close. Then any FTFL rounding algorithm can be applied to solve this FTFL instance, and the solution trivially maps into a solution for the corresponding FTFP instance. Moreover, the approximation ratio for FTFL is preserved for FTFP. Given that FTFL has a 1.7245-approximation algorithm by Byrka, Srinivasan and Swamy [6], it is easy to see that FTFP has an approximation algorithm with the same ratio.

4.1.2 Asymptotic Approximation Ratio for Large Demands

When all demands are large, say $R_{\min} = \omega(|\mathbb{F}|)$, where $R_{\min} \stackrel{\text{def}}{=} \min_{j \in \mathbb{C}} r_j$ is the minimum demand among all clients, then after we apply demand reduction, the integral part has a cost that dominates the overall cost. Since we have a ratio 1 for the integral part, we have a ratio asymptotically close to 1 in the overall cost.

4.2 Adaptive Partition

In this section we develop our second technique, which we call *adaptive partitioning*. Given an FTFP instance and an optimal fractional solution $(\mathbf{x}^*, \mathbf{y}^*)$ to LP (3.1), we split each client j into r_j individual *unit demand points* (or just *demands*), and we split each site i into no more than $|\mathbb{F}| + 2R|\mathbb{C}|^2$ *facility points* (or *facilities*), where $R = \max_{j \in \mathbb{C}} r_j$. We denote the demand set by $\overline{\mathbb{C}}$ and the facility set by $\overline{\mathbb{F}}$, respectively. We will also partition $(\mathbf{x}^*, \mathbf{y}^*)$ into a fractional solution $(\bar{\mathbf{x}}, \bar{\mathbf{y}})$ for the split instance. We will typically use symbols ν and μ to index demands and facilities respectively, that is $\bar{\mathbf{x}} = (\bar{x}_{\mu\nu})$ and $\bar{\mathbf{y}} = (\bar{y}_\mu)$. As before, the *neighborhood of a demand ν* is $\overline{N}(\nu) = \{\mu \in \overline{\mathbb{F}} : \bar{x}_{\mu\nu} > 0\}$. We will use notation $\nu \in j$ to mean that ν is a demand of client j ; similarly, $\mu \in i$ means that facility μ is on site i . Different demands of the same client (that is, $\nu, \nu' \in j$) are called *siblings*. Further, we use the convention that $f_\mu = f_i$ for $\mu \in i$, $\alpha_\nu^* = \alpha_j^*$ for $\nu \in j$ and $d_{\mu\nu} = d_{\mu j} = d_{ij}$ for $\mu \in i$ and $\nu \in j$. We define $C_\nu^{\text{avg}} = \sum_{\mu \in \overline{N}(\nu)} d_{\mu\nu} \bar{x}_{\mu\nu} = \sum_{\mu \in \overline{\mathbb{F}}} d_{\mu\nu} \bar{x}_{\mu\nu}$. One can think of C_ν^{avg} as the average connection cost of demand ν , if we chose a connection to facility μ with probability $\bar{x}_{\mu\nu}$. In our partitioned fractional solution we guarantee for every ν that $\sum_{\mu \in \overline{\mathbb{F}}} \bar{x}_{\mu\nu} = 1$.

Some demands in $\overline{\mathbb{C}}$ will be designated as *primary demands* and the set of primary demands will be denoted by P . By definition we have $P \subseteq \overline{\mathbb{C}}$. In addition, we will use the overlap structure between demand neighborhoods to define a mapping that assigns each demand $\nu \in \overline{\mathbb{C}}$ to some primary demand $\kappa \in P$. As shown in the rounding algorithms in later sections, for each primary demand we guarantee exactly one open facility in its neighborhood, while for a non-primary demand, there is constant probability that none of its neighbors open. In this case we estimate its connection cost by the distance to the facility opened in its assigned primary demand's neighborhood. For this reason the connection cost of a primary demand must be “small” compared to the non-primary demands assigned to it. We also need sibling demands assigned to different primary demands to satisfy the fault-tolerance requirement. Specifically, this partitioning will be constructed to satisfy a number of properties that are detailed below.

(PS) *Partitioned solution.* Vector $(\bar{\mathbf{x}}, \bar{\mathbf{y}})$ is a partition of $(\mathbf{x}^*, \mathbf{y}^*)$, with unit-value demands, that is:

1. $\sum_{\mu \in \overline{\mathbb{F}}} \bar{x}_{\mu\nu} = 1$ for each demand $\nu \in \overline{\mathbb{C}}$.
2. $\sum_{\mu \in i, \nu \in j} \bar{x}_{\mu\nu} = x_{ij}^*$ for each site $i \in \mathbb{F}$ and client $j \in \mathbb{C}$.
3. $\sum_{\mu \in i} \bar{y}_{\mu} = y_i^*$ for each site $i \in \mathbb{F}$.

(CO) *Completeness.* Solution $(\bar{\mathbf{x}}, \bar{\mathbf{y}})$ is complete, that is $\bar{x}_{\mu\nu} \neq 0$ implies $\bar{x}_{\mu\nu} = \bar{y}_{\mu}$, for all $\mu \in \overline{\mathbb{F}}, \nu \in \overline{\mathbb{C}}$.

(PD) *Primary demands.* Primary demands satisfy the following conditions:

1. For any two different primary demands $\kappa, \kappa' \in P$ we have $\overline{N}(\kappa) \cap \overline{N}(\kappa') = \emptyset$.
2. For each site $i \in \mathbb{F}$, $\sum_{\mu \in i} \sum_{\kappa \in P} \bar{x}_{\mu\kappa} \leq y_i^*$.
3. Each demand $\nu \in \overline{\mathbb{C}}$ is assigned to one primary demand $\kappa \in P$ such that
 - (a) $\overline{N}(\nu) \cap \overline{N}(\kappa) \neq \emptyset$, and
 - (b) $C_\nu^{\text{avg}} + \alpha_\nu^* \geq C_\kappa^{\text{avg}} + \alpha_\kappa^*$.

(SI) *Siblings*. For any pair ν, ν' of different siblings we have

1. $\overline{N}(\nu) \cap \overline{N}(\nu') = \emptyset$.
2. If ν is assigned to a primary demand κ then $\overline{N}(\nu') \cap \overline{N}(\kappa) = \emptyset$. In particular, by Property (PD.3(a)), this implies that different sibling demands are assigned to different primary demands.

As we shall demonstrate in later sections, these properties allow us to extend known UFL rounding algorithms to obtain an integral solution to our FTFP problem with a matching approximation ratio. Our partitioning is “adaptive” in the sense that it is constructed one demand at a time, and the connection values for the demands of a client depend on the choice of earlier demands, of this or other clients, and their connection values. We would like to point out that the adaptive partitioning process for the 1.575-approximation algorithm (Section 5.3) is more subtle than that for the 3-approximation (Section 5.1) and the 1.736-approximation algorithms (Section 5.2), due to the introduction of close and far neighborhood.

Implementation of Adaptive Partitioning. We now describe an algorithm for partitioning the instance and the fractional solution so that the properties (PS), (CO), (PD),

and (SI) are satisfied. Recall that $\overline{\mathbb{F}}$ and $\overline{\mathbb{C}}$, respectively, denote the sets of facilities and demands that will be created in this stage, and $(\bar{\mathbf{x}}, \bar{\mathbf{y}})$ is the partitioned solution to be computed.

The adaptive partitioning algorithm consists of two phases: Phase 1 is called the partitioning phase and Phase 2 is called the augmenting phase. Phase 1 is done in iterations, where in each iteration we find the “best” client j and create a new demand ν out of it. This demand either becomes a primary demand itself, or it is assigned to some existing primary demand. We call a client j *exhausted* when all its r_j demands have been created and assigned to some primary demands. Phase 1 completes when all clients are exhausted. In Phase 2 we ensure that every demand has a total connection values $\bar{x}_{\mu\nu}$ equal to 1, that is condition (PS.1).

For each site i we will initially create one “big” facility μ with initial value $\bar{y}_\mu = y_i^*$. While we partition the instance, creating new demands and connections, this facility may end up being split into more facilities to preserve completeness of the fractional solution. Also, we will gradually decrease the fractional connection vector for each client j , to account for the demands already created for j and their connection values. These decreased connection values will be stored in an auxiliary vector $\tilde{\mathbf{x}}$. The intuition is that $\tilde{\mathbf{x}}$ represents the part of \mathbf{x}^* that still has not been allocated to existing demands and future demands can use $\tilde{\mathbf{x}}$ for their connections. For technical reasons, $\tilde{\mathbf{x}}$ will be indexed by facilities (rather than sites) and clients, that is $\tilde{\mathbf{x}} = (\tilde{x}_{\mu j})$. At the beginning, we set $\tilde{x}_{\mu j} \leftarrow x_{ij}^*$ for each $j \in \mathbb{C}$, where $\mu \in i$ is the single facility created initially at site i . At each step, whenever we create a new demand ν for a client j , we will define its values $\bar{x}_{\mu\nu}$ and appropriately reduce the

values $\tilde{x}_{\mu j}$, for all facilities μ . We will deal with two types of neighborhoods, with respect to $\tilde{\mathbf{x}}$ and $\bar{\mathbf{x}}$, that is $\tilde{N}(j) = \{\mu \in \bar{\mathbb{F}} : \tilde{x}_{\mu j} > 0\}$ for $j \in \mathbb{C}$ and $\bar{N}(\nu) = \{\mu \in \bar{\mathbb{F}} : \bar{x}_{\mu\nu} > 0\}$ for $\nu \in \bar{\mathbb{C}}$. During this process we preserve the completeness (CO) of the fractional solutions $\tilde{\mathbf{x}}$ and $\bar{\mathbf{x}}$. More precisely, the following properties will hold for every facility μ after every iteration:

- (c1) For each demand ν either $\bar{x}_{\mu\nu} = 0$ or $\bar{x}_{\mu\nu} = \bar{y}_\mu$. This is the same condition as condition (CO), yet we repeat it here as (c1) needs to hold after every iteration, while condition (CO) only applies to the final partitioned fractional solution $(\bar{\mathbf{x}}, \bar{\mathbf{y}})$.
- (c2) For each client j , either $\tilde{x}_{\mu j} = 0$ or $\tilde{x}_{\mu j} = \bar{y}_\mu$.

A full description of the algorithm is given in Pseudocode 1. Initially, the set U of non-exhausted clients contains all clients, the set $\bar{\mathbb{C}}$ of demands is empty, the set $\bar{\mathbb{F}}$ of facilities consists of one facility μ on each site i with $\bar{y}_\mu = y_i^*$, and the set P of primary demands is empty (Lines 1–4). In one iteration of the while loop (Lines 5–8), for each client j we compute a quantity called $\text{tcc}(j)$ (tentative connection cost), that represents the average distance from j to the set $\tilde{N}_1(j)$ of the nearest facilities μ whose total connection value to j (the sum of $\tilde{x}_{\mu j}$'s) equals 1. This set is computed by Procedure `NEARESTUNITCHUNK()` (see Pseudocode 2, Lines 1–9), which adds facilities to $\tilde{N}_1(j)$ in order of nondecreasing distance, until the total connection value is exactly 1. (The procedure actually uses the \bar{y}_μ values, which are equal to the connection values, by the completeness condition (c2).) This may require splitting the last added facility and adjusting the connection values so that conditions (c1) and (c2) are preserved.

The next step is to pick a client p with minimum $\text{tcc}(p) + \alpha_p^*$ and create a demand ν for p (Lines 9–10). If $\tilde{N}_1(p)$ overlaps the neighborhood of some existing primary demand κ (if there are multiple such κ 's, pick any of them), we assign ν to κ , and ν acquires all the connection values $\tilde{x}_{\mu p}$ between client p and facility μ in $\tilde{N}(p) \cap \overline{N}(\kappa)$ (Lines 11–13). Note that although we check for overlap with $\tilde{N}_1(p)$, we then move all facilities in the intersection with $\tilde{N}(p)$, a bigger set, into $\overline{N}(\nu)$. The other case is when $\tilde{N}_1(p)$ is disjoint from the neighborhoods of all existing primary demands. Then, in Lines 15–16, ν becomes itself a primary demand and we assign ν to itself. It also inherits the connection values to all facilities $\mu \in \tilde{N}_1(p)$ from p (recall that $\tilde{x}_{\mu p} = \bar{y}_\mu$), with all other $\bar{x}_{\mu\nu}$ values set to 0.

At this point all primary demands satisfy Property (PS.1), but this may not be true for non-primary demands. For those demands we still may need to adjust the $\bar{x}_{\mu\nu}$ values so that the total connection value for ν , that is $\text{conn}(\nu) \stackrel{\text{def}}{=} \sum_{\mu \in \mathbb{F}} \bar{x}_{\mu\nu}$, is equal 1. This is accomplished by Procedure AUGMENTTOUNIT() (definition in Pseudocode 2, Lines 10–21) that allocates to $\nu \in j$ some of the remaining connection values $\tilde{x}_{\mu j}$ of client j (Lines 19–21). AUGMENTTOUNIT() will repeatedly pick any facility η with $\tilde{x}_{\eta j} > 0$. If $\tilde{x}_{\eta j} \leq 1 - \text{conn}(\nu)$, then the connection value $\tilde{x}_{\eta j}$ is reassigned to ν . Otherwise, $\tilde{x}_{\eta j} > 1 - \text{conn}(\nu)$, in which case we split η so that connecting ν to one of the created copies of η will make $\text{conn}(\nu)$ equal 1, and we'll be done.

Notice that we start with $|\mathbb{F}|$ facilities and in each iteration of the while loop in Line 5 (Pseudocode 1) each client causes at most one split. We have a total of no more than $R|\mathbb{C}|$ iterations as in each iteration we create one demand. (Recall that $R = \max_j r_j$.) In Phase 2 we do an augment step for each demand ν and this creates no more than $R|\mathbb{C}|$ new

facilities. So the total number of facilities we created will be at most $|\mathbb{F}| + R|\mathbb{C}|^2 + R|\mathbb{C}| \leq |\mathbb{F}| + 2R|\mathbb{C}|^2$, which is polynomial in $|\mathbb{F}| + |\mathbb{C}|$ due to our earlier bound on R .

Example. We now illustrate our partitioning algorithm with an example, where the FTFP instance has four sites and four clients. The demands are $r_1 = 1$ and $r_2 = r_3 = r_4 = 2$. The facility costs are $f_i = 1$ for all i . The distances are defined as follows: $d_{ii} = 3$ for $i = 1, 2, 3, 4$ and $d_{ij} = 1$ for all $i \neq j$. Solving the LP(3.1), we obtain the fractional solution given in Table 4.1a.

x_{ij}^*	1	2	3	4	y_i^*	$\bar{x}_{\mu\nu}$	1'	2'	2''	3'	3''	4'	4''	\bar{y}_μ
1	0	$\frac{4}{3}$	$\frac{4}{3}$	$\frac{4}{3}$	$\frac{4}{3}$	$\dot{1}$	0	1	0	1	0	1	0	1
2	$\frac{1}{3}$	0	$\frac{1}{3}$	$\frac{1}{3}$	$\frac{1}{3}$	$\ddot{1}$	0	0	$\frac{1}{3}$	0	$\frac{1}{3}$	0	$\frac{1}{3}$	$\frac{1}{3}$
3	$\frac{1}{3}$	$\frac{1}{3}$	0	$\frac{1}{3}$	$\frac{1}{3}$	$\dot{2}$	$\frac{1}{3}$	0	0	0	$\frac{1}{3}$	0	$\frac{1}{3}$	$\frac{1}{3}$
4	$\frac{1}{3}$	$\frac{1}{3}$	$\frac{1}{3}$	0	$\frac{1}{3}$	$\dot{3}$	$\frac{1}{3}$	0	$\frac{1}{3}$	0	0	0	$\frac{1}{3}$	$\frac{1}{3}$
						$\dot{4}$	$\frac{1}{3}$	0	$\frac{1}{3}$	0	$\frac{1}{3}$	0	0	$\frac{1}{3}$

(a)

$\bar{x}_{\mu\nu}$	1'	2'	2''	3'	3''	4'	4''	\bar{y}_μ
$\dot{1}$	0	1	0	1	0	1	0	1
$\ddot{1}$	0	0	$\frac{1}{3}$	0	$\frac{1}{3}$	0	$\frac{1}{3}$	$\frac{1}{3}$
$\dot{2}$	$\frac{1}{3}$	0	0	0	$\frac{1}{3}$	0	$\frac{1}{3}$	$\frac{1}{3}$
$\dot{3}$	$\frac{1}{3}$	0	$\frac{1}{3}$	0	0	0	$\frac{1}{3}$	$\frac{1}{3}$
$\dot{4}$	$\frac{1}{3}$	0	$\frac{1}{3}$	0	$\frac{1}{3}$	0	0	$\frac{1}{3}$

(b)

Table 4.1: An example of an execution of the partitioning algorithm. (a) An optimal fractional solution x^*, y^* . (b) The partitioned solution. j' and j'' denote the first and second demand of a client j , and i and \bar{i} denote the first and second facility at site i .

It is easily seen that the fractional solution in Table 4.1a is optimal and complete ($x_{ij}^* > 0$ implies $x_{ij}^* = y_i^*$). The dual optimal solution has all $\alpha_j^* = 4/3$ for $j = 1, 2, 3, 4$.

Now we perform Phase 1, the adaptive partitioning, following the description in

Pseudocode 1. To streamline the presentation, we assume that all ties are broken in favor of lower-numbered clients, demands or facilities. First we create one facility at each of the four sites, denoted as $\dot{1}$, $\dot{2}$, $\dot{3}$ and $\dot{4}$ (Line 2–4, Pseudocode 1). We then execute the “while” loop in Line 5 Pseudocode 1. This loop will have seven iterations. Consider the first iteration. In Line 7–8 we compute $\text{tcc}(j)$ for each client $j = 1, 2, 3, 4$ in U . When computing $\tilde{N}_1(2)$, facility $\dot{1}$ will get split into $\dot{1}$ and $\ddot{1}$ with $\bar{y}_1 = 1$ and $\bar{y}_1 = 1/3$. (This will happen in Line 4–7 of Pseudocode 2.) Then, in Line 9 we will pick client $p = 1$ and create a demand denoted as $1'$ (see Table 4.1b). Since there are no primary demands yet, we make $1'$ a primary demand with $\bar{N}(1') = \tilde{N}_1(1) = \{\dot{2}, \dot{3}, \dot{4}\}$. Notice that client 1 is exhausted after this iteration and U becomes $\{2, 3, 4\}$.

In the second iteration we compute $\text{tcc}(j)$ for $j = 2, 3, 4$ and pick client $p = 2$, from which we create a new demand $2'$. We have $\tilde{N}_1(2) = \{\dot{1}\}$, which is disjoint from $\bar{N}(1')$. So we create a demand $2'$ and make it primary, and set $\bar{N}(2') = \{\dot{1}\}$. In the third iteration we compute $\text{tcc}(j)$ for $j = 2, 3, 4$ and again we pick client $p = 2$. Since $\tilde{N}_1(2) = \{\ddot{1}, \dot{3}, \dot{4}\}$ overlaps with $\bar{N}(1')$, we create a demand $2''$ and assign it to $1'$. We also set $\bar{N}(2'') = \bar{N}(1') \cap \tilde{N}(2) = \{\dot{3}, \dot{4}\}$. After this iteration client 2 is exhausted and we have $U = \{3, 4\}$.

In the fourth iteration we compute $\text{tcc}(j)$ for client $j = 3, 4$. We pick $p = 3$ and create demand $3'$. Since $\tilde{N}_1(3) = \{\dot{1}\}$ overlaps $\bar{N}(2')$, we assign $3'$ to $2'$ and set $\bar{N}(3') = \{\dot{1}\}$. In the fifth iteration we compute $\text{tcc}(j)$ for client $j = 3, 4$ and pick $p = 3$ again. At this time $\tilde{N}_1(3) = \{\ddot{1}, \dot{2}, \dot{4}\}$, which overlaps with $\bar{N}(1')$. So we create a demand $3''$ and assign it to $1'$, as well as set $\bar{N}(3'') = \{\dot{2}, \dot{4}\}$.

In the last two iterations we will pick client $p = 4$ twice and create demands $4'$ and $4''$. For $4'$ we have $\tilde{N}_1(4) = \{\dot{1}\}$ so we assign $4'$ to $2'$ and set $\bar{N}(4') = \{\dot{1}\}$. For $4''$ we have $\tilde{N}_1(4) = \{\ddot{1}, \dot{2}, \dot{3}\}$ and we assign it to $1'$, as well as set $\bar{N}(4'') = \{\dot{2}, \dot{3}\}$.

Now that all clients are exhausted we perform Phase 2, the augmenting phase, to construct a fractional solution in which all demands have total connection value equal to 1. We iterate through each of the seven demands created, that is $1', 2', 2'', 3', 3'', 4', 4''$. $1'$ and $2'$ already have neighborhoods with total connection value of 1, so nothing will change in the first two iterations. $2''$ has $\dot{3}, \dot{4}$ in its neighborhood, with total connection value of $2/3$, and $\tilde{N}(2) = \{\ddot{1}\}$ at this time, so we add $\ddot{1}$ into $\bar{N}(2'')$ to make $\bar{N}(2'') = \{\ddot{1}, \dot{3}, \dot{4}\}$ and now $2''$ has total connection value of 1. Similarly, $3''$ and $4''$ each get $\ddot{1}$ added to their neighborhood and end up with total connection value of 1. The other two demands, namely $3'$ and $4'$, each have $\dot{1}$ in its neighborhood so each of them has already its total connection value equal 1. This completes Phase 2.

The final partitioned fractional solution is given in Table 4.1b. We have created a total of five facilities $\dot{1}, \ddot{1}, \dot{2}, \dot{3}, \dot{4}$, and seven demands, $1', 2', 2'', 3', 3'', 4', 4''$. It can be verified that all the stated properties are satisfied.

Correctness. We now show that all the required properties (PS), (CO), (PD) and (SI) are satisfied by the above construction.

Properties (PS) and (CO) follow directly from the algorithm. (CO) is implied by the completeness condition (c1) that the algorithm maintains after each iteration. Condition (PS.1) is a result of calling Procedure AUGMENTTOUNIT() in Line 21. To see that (PS.2) holds, note that at each step the algorithm maintains the invariant that, for ev-

ery $i \in \mathbb{F}$ and $j \in \mathbb{C}$, we have $\sum_{\mu \in i} \sum_{\nu \in j} \bar{x}_{\mu\nu} + \sum_{\mu \in i} \tilde{x}_{\mu j} = x_{ij}^*$. In the end, we will create r_j demands for each client j , with each demand $\nu \in j$ satisfying (PS.1), and thus $\sum_{\nu \in j} \sum_{\mu \in \mathbb{F}} \bar{x}_{\mu\nu} = r_j$. This implies that $\tilde{x}_{\mu j} = 0$ for every facility $\mu \in \mathbb{F}$, and (PS.2) follows. (PS.3) holds because every time we split a facility μ into μ' and μ'' , the sum of $\bar{y}_{\mu'}$ and $\bar{y}_{\mu''}$ is equal to the old value of \bar{y}_{μ} .

Now we deal with properties in group (PD). First, (PD.1) follows directly from the algorithm, Pseudocode 1 (Lines 14–16), since every primary demand has its neighborhood fixed when created, and that neighborhood is disjoint from those of the existing primary demands.

Property (PD.2) follows from (PD.1), (CO) and (PS.3). In more detail, it can be justified as follows. By (PD.1), for each $\mu \in i$ there is at most one $\kappa \in P$ with $\bar{x}_{\mu\kappa} > 0$ and we have $\bar{x}_{\mu\kappa} = \bar{y}_{\mu}$ due to (CO). Let $K \subseteq i$ be the set of those μ 's for which such $\kappa \in P$ exists, and denote this κ by κ_{μ} . Then, using conditions (CO) and (PS.3), we have $\sum_{\mu \in i} \sum_{\kappa \in P} \bar{x}_{\mu\kappa} = \sum_{\mu \in K} \bar{x}_{\mu\kappa_{\mu}} = \sum_{\mu \in K} \bar{y}_{\mu} \leq \sum_{\mu \in i} \bar{y}_{\mu} = y_i^*$.

Property (PD.3(a)) follows from the way the algorithm assigns primary demands. When demand ν of client p is assigned to a primary demand κ in Lines 11–13 of Pseudocode 1, we move all facilities in $\tilde{N}(p) \cap \bar{N}(\kappa)$ (the intersection is nonempty) into $\bar{N}(\nu)$, and we never remove a facility from $\bar{N}(\nu)$. We postpone the proof for (PD.3(b)) to Lemma 12.

Finally we argue that the properties in group (SI) hold. (SI.1) is easy, since for any client j , each facility μ is added to the neighborhood of at most one demand $\nu \in j$, by setting $\bar{x}_{\mu\nu}$ to \bar{y}_{μ} , while other siblings ν' of ν have $\bar{x}_{\mu\nu'} = 0$. Note that right after a demand $\nu \in p$ is created, its neighborhood is disjoint from the neighborhood of p , that is

$\overline{N}(\nu) \cap \tilde{N}(p) = \emptyset$, by Lines 11–13 of the algorithm. Thus all demands of p created later will have neighborhoods disjoint from the set $\overline{N}(\nu)$ before the augmenting phase 2. Furthermore, Procedure AUGMENTTOUNIT() preserves this property, because when it adds a facility to $\overline{N}(\nu)$ then it removes it from $\tilde{N}(p)$, and in case of splitting, one resulting facility is added to $\overline{N}(\nu)$ and the other to $\tilde{N}(p)$. Property (SI.2) is shown below in Lemma 10.

It remains to show Properties (PD.3(b)) and (SI.2). We show them in the lemmas below, thus completing the description of our adaptive partition process.

Lemma 10 *Property (SI.2) holds after the Adaptive Partitioning stage.*

Proof. Let ν_1, \dots, ν_{r_j} be the demands of a client $j \in \mathbb{C}$, listed in the order of creation, and, for each $q = 1, 2, \dots, r_j$, denote by κ_q the primary demand that ν_q is assigned to. After the completion of Phase 1 of Pseudocode 1 (Lines 5–18), we have $\overline{N}(\nu_s) \subseteq \overline{N}(\kappa_s)$ for $s = 1, \dots, r_j$. Since any two primary demands have disjoint neighborhoods, we have $\overline{N}(\nu_s) \cap \overline{N}(\kappa_q) = \emptyset$ for any $s \neq q$, that is Property (SI.2) holds right after Phase 1.

After Phase 1 all neighborhoods $\overline{N}(\kappa_s)$, $s = 1, \dots, r_j$ have already been fixed and they do not change in Phase 2. None of the facilities in $\tilde{N}(j)$ appear in any of $\overline{N}(\kappa_s)$ for $s = 1, \dots, r_j$, by the way we allocate facilities in Lines 13 and 16. Therefore during the augmentation process in Phase 2, when we add facilities from $\tilde{N}(j)$ to $\overline{N}(\nu)$, for some $\nu \in j$ (Line 19–21 of Pseudocode 1), all the required disjointness conditions will be preserved. ■

We need one more lemma before proving our last property (PD.3(b)). For a client j and a demand ν , we use notation $\text{tcc}^\nu(j)$ for the value of $\text{tcc}(j)$ at the time when ν was created. (It is not necessary that $\nu \in j$ but we assume that j is not exhausted at that time.)

Lemma 11 *Let η and ν be two demands, with η created no later than ν , and let $j \in \mathbb{C}$ be a client that is not exhausted when ν is created. Then we have*

$$(a) \text{ tcc}^\eta(j) \leq \text{tcc}^\nu(j), \text{ and}$$

$$(b) \text{ if } \nu \in j \text{ then } \text{tcc}^\eta(j) \leq C_\nu^{\text{avg}}.$$

Proof. We focus first on the time when demand η is about to be created, right after the call to `NEARESTUNITCHUNK()` in Pseudocode 1, Line 7. Let $\tilde{N}(j) = \{\mu_1, \dots, \mu_q\}$ with all facilities μ_s ordered according to nondecreasing distance from j . Consider the following linear program:

$$\begin{aligned} & \text{minimize} && \sum_s d_{\mu_s j} z_s \\ & \text{subject to} && \sum_s z_s \geq 1 \\ & && 0 \leq z_s \leq \tilde{x}_{\mu_s j} \quad \text{for all } s \end{aligned}$$

This is a fractional minimum knapsack covering problem (with knapsack size equal 1) and its optimal fractional solution is the greedy solution, whose value is exactly $\text{tcc}^\eta(j)$.

On the other hand, we claim that $\text{tcc}^\nu(j)$ can be thought of as the value of some feasible solution to this linear program, and that the same is true for C_ν^{avg} if $\nu \in j$. Indeed, each of these quantities involves some later values $\tilde{x}_{\mu j}$, where μ could be one of the facilities μ_s or a new facility obtained from splitting. For each s , however, the sum of all values $\tilde{x}_{\mu j}$, over the facilities μ that were split from μ_s , cannot exceed the value $\tilde{x}_{\mu_s j}$ at the time when η was created, because splitting facilities preserves this sum and creating new demands for j can only decrease it. Therefore both quantities $\text{tcc}^\nu(j)$ and C_ν^{avg} (for $\nu \in j$) correspond

to some choice of the z_s variables (adding up to 1), and the lemma follows. ■

Lemma 12 *Property (PD.3(b)) holds after the Adaptive Partitioning stage.*

Proof. Suppose that demand $\nu \in j$ is assigned to some primary demand $\kappa \in p$.

Then

$$C_{\kappa}^{\text{avg}} + \alpha_{\kappa}^* = \text{tcc}^{\kappa}(p) + \alpha_p^* \leq \text{tcc}^{\kappa}(j) + \alpha_j^* \leq C_{\nu}^{\text{avg}} + \alpha_{\nu}^*.$$

We now justify this derivation. By definition we have $\alpha_{\kappa}^* = \alpha_p^*$. Further, by the algorithm, if κ is a primary demand of client p , then C_{κ}^{avg} is equal to $\text{tcc}(p)$ computed when κ is created, which is exactly $\text{tcc}^{\kappa}(p)$. Thus the first equation is true. The first inequality follows from the choice of p in Line 9 in Pseudocode 1. The last inequality holds because $\alpha_j^* = \alpha_{\nu}^*$ (due to $\nu \in j$), and because $\text{tcc}^{\kappa}(j) \leq C_{\nu}^{\text{avg}}$, which follows from Lemma 11. ■

We have thus proved that all properties (PS), (CO), (PD) and (SI) hold for our partitioned fractional solution $(\bar{\mathbf{x}}, \bar{\mathbf{y}})$. In the following sections we show how to use these properties to round the fractional solution to an approximate integral solution. For the 3-approximation algorithm (Section 5.1) and the 1.736-approximation algorithm (Section 5.2), the first phase of the algorithm is exactly the same partition process as described above. However, the 1.575-approximation algorithm (Section 5.3) demands a more sophisticated partitioning process as the interplay between close and far neighborhood of sibling demands result in more delicate properties that our partitioned fractional solution must satisfy.

Pseudocode 1 Algorithm: Adaptive Partitioning

Input: $\mathbb{F}, \mathbb{C}, (\mathbf{x}^*, \mathbf{y}^*)$ **Output:** $\overline{\mathbb{F}}, \overline{\mathbb{C}}, (\bar{\mathbf{x}}, \bar{\mathbf{y}})$ \triangleright Unspecified $\bar{x}_{\mu\nu}$'s and $\tilde{x}_{\mu j}$'s are assumed to be 01: $\tilde{\mathbf{r}} \leftarrow \mathbf{r}, U \leftarrow \mathbb{C}, \overline{\mathbb{F}} \leftarrow \emptyset, \overline{\mathbb{C}} \leftarrow \emptyset, P \leftarrow \emptyset$ \triangleright Phase 12: **for** each site $i \in \mathbb{F}$ **do**3: create a facility μ at i and add μ to $\overline{\mathbb{F}}$ 4: $\bar{y}_\mu \leftarrow y_i^*$ and $\tilde{x}_{\mu j} \leftarrow x_{ij}^*$ for each $j \in \mathbb{C}$ 5: **while** $U \neq \emptyset$ **do**6: **for** each $j \in U$ **do**7: $\tilde{N}_1(j) \leftarrow \text{NEARESTUNITCHUNK}(j, \overline{\mathbb{F}}, \tilde{\mathbf{x}}, \bar{\mathbf{x}}, \bar{\mathbf{y}})$ \triangleright see Pseudocode 28: $\text{tcc}(j) \leftarrow \sum_{\mu \in \tilde{N}_1(j)} d_{\mu j} \cdot \tilde{x}_{\mu j}$ 9: $p \leftarrow \arg \min_{j \in U} \{\text{tcc}(j) + \alpha_j^*\}$ 10: create a new demand ν for client p 11: **if** $\tilde{N}_1(p) \cap \overline{N}(\kappa) \neq \emptyset$ for some primary demand $\kappa \in P$ **then**12: assign ν to κ 13: $\bar{x}_{\mu\nu} \leftarrow \tilde{x}_{\mu p}$ and $\tilde{x}_{\mu p} \leftarrow 0$ for each $\mu \in \tilde{N}_1(p) \cap \overline{N}(\kappa)$ 14: **else**15: make ν primary, $P \leftarrow P \cup \{\nu\}$, assign ν to itself16: set $\bar{x}_{\mu\nu} \leftarrow \tilde{x}_{\mu p}$ and $\tilde{x}_{\mu p} \leftarrow 0$ for each $\mu \in \tilde{N}_1(p)$ 17: $\overline{\mathbb{C}} \leftarrow \overline{\mathbb{C}} \cup \{\nu\}, \tilde{r}_p \leftarrow \tilde{r}_p - 1$ 18: **if** $\tilde{r}_p = 0$ **then** $U \leftarrow U \setminus \{p\}$ 19: **for** each client $j \in \mathbb{C}$ **do** \triangleright Phase 220: **for** each demand $\nu \in j$ **do** \triangleright each client j has r_j demands21: **if** $\sum_{\mu \in \overline{N}(\nu)} \bar{x}_{\mu\nu} < 1$ **then** $\text{AUGMENTTOUNIT}(\nu, j, \overline{\mathbb{F}}, \tilde{\mathbf{x}}, \bar{\mathbf{x}}, \bar{\mathbf{y}})$ \triangleright see Pseudocode 2

Pseudocode 2 Helper functions used in Pseudocode 1

```

1: function NEARESTUNITCHUNK( $j, \bar{\mathbb{F}}, \tilde{\mathbf{x}}, \bar{\mathbf{x}}, \bar{\mathbf{y}}$ )           ▷ upon return,  $\sum_{\mu \in \tilde{N}_1(j)} \tilde{x}_{\mu j} = 1$ 

2:   Let  $\tilde{N}(j) = \{\mu_1, \dots, \mu_q\}$  where  $d_{\mu_1 j} \leq d_{\mu_2 j} \leq \dots \leq d_{\mu_q j}$ 

3:   Let  $l$  be such that  $\sum_{k=1}^l \bar{y}_{\mu_k} \geq 1$  and  $\sum_{k=1}^{l-1} \bar{y}_{\mu_k} < 1$ 

4:   Create a new facility  $\sigma$  at the same site as  $\mu_l$  and add it to  $\bar{\mathbb{F}}$            ▷ split  $\mu_l$ 

5:   Set  $\bar{y}_\sigma \leftarrow \sum_{k=1}^l \bar{y}_{\mu_k} - 1$  and  $\bar{y}_{\mu_l} \leftarrow \bar{y}_{\mu_l} - \bar{y}_\sigma$ 

6:   For each  $\nu \in \bar{\mathbb{C}}$  with  $\bar{x}_{\mu_l \nu} > 0$  set  $\bar{x}_{\mu_l \nu} \leftarrow \bar{y}_{\mu_l}$  and  $\bar{x}_{\sigma \nu} \leftarrow \bar{y}_\sigma$ 

7:   For each  $j' \in \bar{\mathbb{C}}$  with  $\tilde{x}_{\mu_l j'} > 0$  (including  $j$ ) set  $\tilde{x}_{\mu_l j'} \leftarrow \bar{y}_{\mu_l}$  and  $\tilde{x}_{\sigma j'} \leftarrow \bar{y}_\sigma$ 

8:   (All other new connection values are set to 0)

9:   return  $\tilde{N}_1(j) = \{\mu_1, \dots, \mu_{l-1}, \mu_l\}$ 

10: function AUGMENTTOUNIT( $\nu, j, \bar{\mathbb{F}}, \tilde{\mathbf{x}}, \bar{\mathbf{x}}, \bar{\mathbf{y}}$ )           ▷  $\nu$  is a demand of client  $j$ 

11:   while  $\sum_{\mu \in \bar{\mathbb{F}}} \bar{x}_{\mu \nu} < 1$  do           ▷ upon return,  $\sum_{\mu \in \bar{N}(\nu)} \bar{x}_{\mu \nu} = 1$ 

12:     Let  $\eta$  be any facility such that  $\tilde{x}_{\eta j} > 0$ 

13:     if  $1 - \sum_{\mu \in \bar{\mathbb{F}}} \bar{x}_{\mu \nu} \geq \tilde{x}_{\eta j}$  then

14:        $\bar{x}_{\eta \nu} \leftarrow \tilde{x}_{\eta j}, \tilde{x}_{\eta j} \leftarrow 0$ 

15:     else

16:       Create a new facility  $\sigma$  at the same site as  $\eta$  and add it to  $\bar{\mathbb{F}}$            ▷ split  $\eta$ 

17:       Let  $\bar{y}_\sigma \leftarrow 1 - \sum_{\mu \in \bar{\mathbb{F}}} \bar{x}_{\mu \nu}, \bar{y}_\eta \leftarrow \bar{y}_\eta - \bar{y}_\sigma$ 

18:       Set  $\bar{x}_{\sigma \nu} \leftarrow \bar{y}_\sigma, \bar{x}_{\eta \nu} \leftarrow 0, \tilde{x}_{\eta j} \leftarrow \bar{y}_\eta, \tilde{x}_{\sigma j} \leftarrow 0$ 

19:       For each  $\nu' \neq \nu$  with  $\bar{x}_{\eta \nu'} > 0$ , set  $\bar{x}_{\eta \nu'} \leftarrow \bar{y}_\eta, \bar{x}_{\sigma \nu'} \leftarrow \bar{y}_\sigma$ 

20:       For each  $j' \neq j$  with  $\tilde{x}_{\eta j'} > 0$ , set  $\tilde{x}_{\eta j'} \leftarrow \bar{y}_\eta, \tilde{x}_{\sigma j'} \leftarrow \bar{y}_\sigma$ 

21:       (All other new connection values are set to 0)

```

Chapter 5

LP-rounding Algorithms

5.1 Algorithm EGUP with Ratio 3

With the partitioned FTFP instance and its associated fractional solution in place, we now begin to introduce our rounding algorithms. The algorithm we describe in this section achieves ratio 3. Although this is still quite far from our best ratio 1.575 that we derive later, we include this algorithm in the paper to illustrate, in a relatively simple setting, how the properties of our partitioned fractional solution are used in rounding it to an integral solution with cost not too far away from an optimal solution. The rounding approach we use here is an extension of the corresponding method for UFL described in [14].

Algorithm EGUP. At a high level, we would open exactly one facility for each primary demand κ , and each non-primary demand is connected to the facility opened for the primary demand it was assigned to.

More precisely, we apply a rounding process, guided by the fractional values (\bar{y}_μ)

and $(\bar{x}_{\mu\nu})$, that produces an integral solution. This integral solution is obtained by choosing a subset of facilities in $\bar{\mathbb{F}}$ to open, and for each demand in $\bar{\mathbb{C}}$, specifying an open facility that this demand will be connected to. For each primary demand $\kappa \in P$, we want to open one facility $\phi(\kappa) \in \bar{N}(\kappa)$. To this end, we use randomization: for each $\mu \in \bar{N}(\kappa)$, we choose $\phi(\kappa) = \mu$ with probability $\bar{x}_{\mu\kappa}$, ensuring that exactly one $\mu \in \bar{N}(\kappa)$ is chosen. Note that $\sum_{\mu \in \bar{N}(\kappa)} \bar{x}_{\mu\kappa} = 1$, so this distribution is well-defined. We open this facility $\phi(\kappa)$ and connect to $\phi(\kappa)$ all demands that are assigned to κ .

In our description above, the algorithm is presented as a randomized algorithm. It can be de-randomized using the method of conditional expectations, which is commonly used in approximation algorithms for facility location problems and standard enough that presenting it here would be redundant. Readers less familiar with this field are recommended to consult [8], where the method of conditional expectations is applied in a context very similar to ours.

Analysis. We now bound the expected facility cost and connection cost by establishing the two lemmas below.

Lemma 13 *The expectation of facility cost F_{EGUP} of our solution is at most F^* .*

Proof. By Property (PD.1), the neighborhoods of primary demands are disjoint. Also, for any primary demand $\kappa \in P$, the probability that a facility $\mu \in \bar{N}(\kappa)$ is chosen as

the open facility $\phi(\kappa)$ is $\bar{x}_{\mu\kappa}$. Hence the expected total facility cost is

$$\begin{aligned}
\mathbb{E}[F_{\text{EGUP}}] &= \sum_{\kappa \in P} \sum_{\mu \in \bar{N}(\kappa)} f_{\mu} \bar{x}_{\mu\kappa} \\
&= \sum_{\kappa \in P} \sum_{\mu \in \bar{\mathbb{F}}} f_{\mu} \bar{x}_{\mu\kappa} \\
&= \sum_{i \in \mathbb{F}} f_i \sum_{\mu \in i} \sum_{\kappa \in P} \bar{x}_{\mu\kappa} \\
&\leq \sum_{i \in \mathbb{F}} f_i y_i^* = F^*,
\end{aligned}$$

where the inequality follows from Property (PD.2). ■

Lemma 14 *The expectation of connection cost C_{EGUP} of our solution is at most $C^* + 2 \cdot \text{LP}^*$.*

Proof. For a primary demand κ , its expected connection cost is C_{κ}^{avg} because we choose facility μ with probability $\bar{x}_{\mu\kappa}$.

Consider a non-primary demand ν assigned to a primary demand $\kappa \in P$. Let μ be any facility in $\bar{N}(\nu) \cap \bar{N}(\kappa)$. Since μ is in both $\bar{N}(\nu)$ and $\bar{N}(\kappa)$, we have $d_{\mu\nu} \leq \alpha_{\nu}^*$ and $d_{\mu\kappa} \leq \alpha_{\kappa}^*$ (This follows from the complementary slackness conditions since $\alpha_{\nu}^* = \beta_{\mu\nu}^* + d_{\mu\nu}$ for each $\mu \in \bar{N}(\nu)$). Thus, applying the triangle inequality, for any fixed choice of facility $\phi(\kappa)$ we have

$$d_{\phi(\kappa)\nu} \leq d_{\phi(\kappa)\kappa} + d_{\mu\kappa} + d_{\mu\nu} \leq d_{\phi(\kappa)\kappa} + \alpha_{\kappa}^* + \alpha_{\nu}^*.$$

Therefore the expected distance from ν to its facility $\phi(\kappa)$ is

$$\begin{aligned}
\mathbb{E}[d_{\phi(\kappa)\nu}] &\leq C_{\kappa}^{\text{avg}} + \alpha_{\kappa}^* + \alpha_{\nu}^* \\
&\leq C_{\nu}^{\text{avg}} + \alpha_{\nu}^* + \alpha_{\nu}^* = C_{\nu}^{\text{avg}} + 2\alpha_{\nu}^*,
\end{aligned}$$

where the second inequality follows from Property (PD.3(b)). From the definition of C_{ν}^{avg}

and Property (PS.2), for any $j \in \mathbb{C}$ we have

$$\begin{aligned}
\sum_{\nu \in j} C_{\nu}^{\text{avg}} &= \sum_{\nu \in j} \sum_{\mu \in \overline{\mathbb{F}}} d_{\mu\nu} \bar{x}_{\mu\nu} \\
&= \sum_{i \in \overline{\mathbb{F}}} d_{ij} \sum_{\nu \in j} \sum_{\mu \in i} \bar{x}_{\mu\nu} \\
&= \sum_{i \in \overline{\mathbb{F}}} d_{ij} x_{ij}^* = C_j^*.
\end{aligned}$$

Thus, summing over all demands, the expected total connection cost is

$$\begin{aligned}
\mathbb{E}[C_{\text{EGUP}}] &\leq \sum_{j \in \mathbb{C}} \sum_{\nu \in j} (C_{\nu}^{\text{avg}} + 2\alpha_{\nu}^*) \\
&= \sum_{j \in \mathbb{C}} (C_j^* + 2r_j\alpha_j^*) = C^* + 2 \cdot \text{LP}^*,
\end{aligned}$$

completing the proof of the lemma. ■

Theorem 15 *Algorithm EGUP is a 3-approximation algorithm.*

Proof. By Property (SI.2), different demands from the same client are assigned to different primary demands, and by (PD.1) each primary demand opens a different facility. This ensures that our solution is feasible, namely each client j is connected to r_j different facilities (some possibly located on the same site). As for the total cost, Lemma 13 and Lemma 14 imply that the total cost is at most $F^* + C^* + 2 \cdot \text{LP}^* = 3 \cdot \text{LP}^* \leq 3 \cdot \text{OPT}$. ■

5.2 Algorithm ECHS with Ratio 1.736

In this section we improve the approximation ratio to $1 + 2/e \approx 1.736$. The improvement comes from a slightly modified rounding process and refined analysis. Note that

the facility opening cost of Algorithm EGUP does not exceed that of the fractional optimum solution, while the connection cost could be far from the optimum, since we connect a non-primary demand to a facility in the neighborhood of its assigned primary demand and then estimate the distance using the triangle inequality. The basic idea to improve the estimate of the connection cost, following the approach of Chudak and Shmoys [8], is to connect each non-primary demand to its nearest neighbor when one is available, and to only use the facility opened by its assigned primary demand when none of its neighbors is open.

Algorithm ECHS. As before, the algorithm starts by solving the linear program and applying the adaptive partitioning algorithm described in Section 4.2 to obtain a partitioned solution (\bar{x}, \bar{y}) . Then we apply the rounding process to compute an integral solution (see Pseudocode 3).

We start, as before, by opening exactly one facility $\phi(\kappa)$ in the neighborhood of each primary demand κ (Line 2). For any non-primary demand ν assigned to κ , we refer to $\phi(\kappa)$ as the *target* facility of ν . In Algorithm EGUP, ν was connected to $\phi(\kappa)$, but in Algorithm ECHS we may be able to find an open facility in ν 's neighborhood and connect ν to this facility. Specifically, the two changes in the algorithm are as follows:

- (1) Each facility μ that is not in the neighborhood of any primary demand is opened, independently, with probability \bar{y}_μ (Lines 4–5). Notice that if $\bar{y}_\mu > 0$ then, due to completeness of the partitioned fractional solution, we have $\bar{y}_\mu = \bar{x}_{\mu\nu}$ for some demand ν . This implies that $\bar{y}_\mu \leq 1$, because $\bar{x}_{\mu\nu} \leq 1$, by (PS.1).

- (2) When connecting demands to facilities, a primary demand κ is connected to the only facility $\phi(\kappa)$ opened in its neighborhood, as before (Line 3). For a non-primary demand ν , if its neighborhood $\overline{N}(\nu)$ has an open facility, we connect ν to the closest open facility in $\overline{N}(\nu)$ (Line 8). Otherwise, we connect ν to its target facility (Line 10).

Pseudocode 3 Algorithm ECHS: Constructing Integral Solution

```

1: for each  $\kappa \in P$  do
2:   choose one  $\phi(\kappa) \in \overline{N}(\kappa)$ , with each  $\mu \in \overline{N}(\kappa)$  chosen as  $\phi(\kappa)$  with probability  $\bar{y}_\mu$ 
3:   open  $\phi(\kappa)$  and connect  $\kappa$  to  $\phi(\kappa)$ 
4: for each  $\mu \in \overline{\mathbb{F}} - \bigcup_{\kappa \in P} \overline{N}(\kappa)$  do
5:   open  $\mu$  with probability  $\bar{y}_\mu$  (independently)
6: for each non-primary demand  $\nu \in \overline{\mathbb{C}}$  do
7:   if any facility in  $\overline{N}(\nu)$  is open then
8:     connect  $\nu$  to the nearest open facility in  $\overline{N}(\nu)$ 
9:   else
10:    connect  $\nu$  to  $\phi(\kappa)$  where  $\kappa$  is  $\nu$ 's assigned primary demand

```

Analysis. We shall first argue that the integral solution thus constructed is feasible, and then we bound the total cost of the solution. Regarding feasibility, the only constraint that is not explicitly enforced by the algorithm is the fault-tolerance requirement; namely that each client j is connected to r_j different facilities. Let ν and ν' be two different sibling demands of client j and let their assigned primary demands be κ and κ' respectively. Due to (SI.2) we know $\kappa \neq \kappa'$. From (SI.1) we have $\overline{N}(\nu) \cap \overline{N}(\nu') = \emptyset$. From (SI.2), we have $\overline{N}(\nu) \cap \overline{N}(\kappa') = \emptyset$ and $\overline{N}(\nu') \cap \overline{N}(\kappa) = \emptyset$. From (PD.1) we have $\overline{N}(\kappa) \cap \overline{N}(\kappa') = \emptyset$. It

follows that $(\overline{N}(\nu) \cup \overline{N}(\kappa)) \cap (\overline{N}(\nu') \cup \overline{N}(\kappa')) = \emptyset$. Since the algorithm connects ν to some facility in $\overline{N}(\nu) \cup \overline{N}(\kappa)$ and ν' to some facility in $\overline{N}(\nu') \cup \overline{N}(\kappa')$, ν and ν' will be connected to different facilities.

We now show that the expected cost of the computed solution is bounded by $(1 + 2/e) \cdot \text{LP}^*$. By (PD.1), every facility may appear in at most one primary demand's neighborhood, and the facilities open in Line 4–5 of Pseudocode 3 do not appear in any primary demand's neighborhood. Therefore, by linearity of expectation, the expected facility cost of Algorithm ECHS is

$$\mathbb{E}[F_{\text{ECHS}}] = \sum_{\mu \in \overline{\mathbb{F}}} f_{\mu} \bar{y}_{\mu} = \sum_{i \in \mathbb{F}} f_i \sum_{\mu \in i} \bar{y}_{\mu} = \sum_{i \in \mathbb{F}} f_i y_i^* = F^*,$$

where the third equality follows from (PS.3).

To bound the connection cost, we adapt an argument of Chudak and Shmoys [8]. Consider a demand ν and denote by C_{ν} the random variable representing the connection cost for ν . Our goal now is to estimate $\mathbb{E}[C_{\nu}]$, the expected value of C_{ν} . Demand ν can either get connected directly to some facility in $\overline{N}(\nu)$ or indirectly to its target facility $\phi(\kappa) \in \overline{N}(\kappa)$, where κ is the primary demand to which ν is assigned. We will analyze these two cases separately.

In our analysis, in this section and the next one, we will use notation

$$D(A, \sigma) = \sum_{\mu \in A} d_{\mu\sigma} \bar{y}_{\mu} / \sum_{\mu \in A} \bar{y}_{\mu}$$

for the average distance between a demand σ and a set A of facilities. Note that, in particular, we have $C_{\nu}^{\text{avg}} = D(\overline{N}(\nu), \nu)$.

We first estimate the expected cost $d_{\phi(\kappa)\nu}$ of the indirect connection. Let Λ^ν denote the event that some facility in $\overline{N}(\nu)$ is opened. Then

$$\mathbb{E}[C_\nu \mid \neg\Lambda^\nu] = \mathbb{E}[d_{\phi(\kappa)\nu} \mid \neg\Lambda^\nu] = D(\overline{N}(\kappa) \setminus \overline{N}(\nu), \nu). \quad (5.1)$$

Note that $\neg\Lambda^\nu$ implies that $\overline{N}(\kappa) \setminus \overline{N}(\nu) \neq \emptyset$, since $\overline{N}(\kappa)$ contains exactly one open facility, namely $\phi(\kappa)$.

Lemma 16 *Let ν be a demand assigned to a primary demand κ , and assume that $\overline{N}(\kappa) \setminus \overline{N}(\nu) \neq \emptyset$. Then*

$$\mathbb{E}[C_\nu \mid \neg\Lambda^\nu] \leq C_\nu^{\text{avg}} + 2\alpha_\nu^*.$$

Proof. By (5.1), we need to show that $D(\overline{N}(\kappa) \setminus \overline{N}(\nu), \nu) \leq C_\nu^{\text{avg}} + 2\alpha_\nu^*$. There are two cases to consider.

Case 1: There exists some $\mu' \in \overline{N}(\kappa) \cap \overline{N}(\nu)$ such that $d_{\mu'\kappa} \leq C_\kappa^{\text{avg}}$. In this case, for every

$\mu \in \overline{N}(\kappa) \setminus \overline{N}(\nu)$, we have

$$d_{\mu\nu} \leq d_{\mu\kappa} + d_{\mu'\kappa} + d_{\mu'\nu} \leq \alpha_\kappa^* + C_\kappa^{\text{avg}} + \alpha_\nu^* \leq C_\nu^{\text{avg}} + 2\alpha_\nu^*,$$

using the triangle inequality, complementary slackness, and (PD.3(b)). By summing

over all $\mu \in \overline{N}(\kappa) \setminus \overline{N}(\nu)$, it follows that $D(\overline{N}(\kappa) \setminus \overline{N}(\nu), \nu) \leq C_\nu^{\text{avg}} + 2\alpha_\nu^*$.

Case 2: Every $\mu' \in \overline{N}(\kappa) \cap \overline{N}(\nu)$ has $d_{\mu'\kappa} > C_\kappa^{\text{avg}}$. Since $C_\kappa^{\text{avg}} = D(\overline{N}(\kappa), \kappa)$, this implies

that $D(\overline{N}(\kappa) \setminus \overline{N}(\nu), \kappa) \leq C_\kappa^{\text{avg}}$. Therefore, choosing an arbitrary $\mu' \in \overline{N}(\kappa) \cap \overline{N}(\nu)$,

we obtain

$$D(\overline{N}(\kappa) \setminus \overline{N}(\nu), \nu) \leq D(\overline{N}(\kappa) \setminus \overline{N}(\nu), \kappa) + d_{\mu'\kappa} + d_{\mu'\nu} \leq C_\kappa^{\text{avg}} + \alpha_\kappa^* + \alpha_\nu^* \leq C_\nu^{\text{avg}} + 2\alpha_\nu^*,$$

where we again use the triangle inequality, complementary slackness, and (PD.3(b)).

Since the lemma holds in both cases, the proof is now complete. ■

We now continue our estimation of the connection cost. The next step of our analysis is to show that

$$\mathbb{E}[C_\nu] \leq C_\nu^{\text{avg}} + \frac{2}{e}\alpha_\nu^*. \quad (5.2)$$

The argument is divided into three cases. The first, easy case is when ν is a primary demand κ . According to the algorithm (see Pseudocode 3, Line 2), we have $C_\kappa = d_{\mu\kappa}$ with probability \bar{y}_μ , for $\mu \in \bar{N}(\kappa)$. Therefore $\mathbb{E}[C_\kappa] = C_\kappa^{\text{avg}}$, so (5.2) holds.

Next, we consider a non-primary demand ν . Let κ be the primary demand that ν is assigned to. We first deal with the sub-case when $\bar{N}(\kappa) \setminus \bar{N}(\nu) = \emptyset$, which is the same as $\bar{N}(\kappa) \subseteq \bar{N}(\nu)$. Property (CO) implies that $\bar{x}_{\mu\nu} = \bar{y}_\mu = \bar{x}_{\mu\kappa}$ for every $\mu \in \bar{N}(\kappa)$, so we have $\sum_{\mu \in \bar{N}(\kappa)} \bar{x}_{\mu\nu} = \sum_{\mu \in \bar{N}(\kappa)} \bar{x}_{\mu\kappa} = 1$, due to (PS.1). On the other hand, we have $\sum_{\mu \in \bar{N}(\nu)} \bar{x}_{\mu\nu} = 1$, and $\bar{x}_{\mu\nu} > 0$ for all $\mu \in \bar{N}(\nu)$. Therefore $\bar{N}(\kappa) = \bar{N}(\nu)$ and C_ν has exactly the same distribution as C_κ . So this case reduces to the first case, namely we have $\mathbb{E}[C_\nu] = C_\nu^{\text{avg}}$, and (5.2) holds.

The last, and only non-trivial case is when $\bar{N}(\kappa) \setminus \bar{N}(\nu) \neq \emptyset$. We handle this case in the following lemma.

Lemma 17 *Assume that $\bar{N}(\kappa) \setminus \bar{N}(\nu) \neq \emptyset$. Then the expected connection cost of ν , conditioned on the event that at least one of its neighbor opens, satisfies*

$$\mathbb{E}[C_\nu \mid \Lambda^\nu] \leq C_\nu^{\text{avg}}.$$

Proof. The proof is similar to an analogous result in [8, 4]. For the sake of completeness we sketch here a simplified argument, adapted to our terminology and nota-

tion. The idea is to consider a different random process that is easier to analyze and whose expected connection cost is not better than that in the algorithm.

We partition $\overline{N}(\nu)$ into groups G_1, \dots, G_k , where two different facilities μ and μ' are put in the same G_s , where $s \in \{1, \dots, k\}$, if they both belong to the same set $\overline{N}(\kappa)$ for some primary demand κ . If some μ is not a neighbor of any primary demand, then it constitutes a singleton group. For each s , let $\bar{d}_s = D(G_s, \nu)$ be the average distance from ν to G_s . Assume that G_1, \dots, G_k are ordered by nondecreasing average distance to ν , that is $\bar{d}_1 \leq \bar{d}_2 \leq \dots \leq \bar{d}_k$. For each group G_s , we select it, independently, with probability $g_s = \sum_{\mu \in G_s} \bar{y}_\mu$. For each selected group G_s , we open exactly one facility in G_s , where each $\mu \in G_s$ is opened with probability $\bar{y}_\mu / \sum_{\eta \in G_s} \bar{y}_\eta$.

So far, this process is the same as that in the algorithm (if restricted to $\overline{N}(\nu)$). However, we connect ν in a slightly different way, by choosing the smallest s for which G_s was selected and connecting ν to the open facility in G_s . This can only increase our expected connection cost, assuming that at least one facility in $\overline{N}(\nu)$ opens, so

$$\begin{aligned} \mathbb{E}[C_\nu \mid \Lambda^\nu] &\leq \frac{1}{\mathbb{P}[\Lambda^\nu]} (\bar{d}_1 g_1 + \bar{d}_2 g_2 (1 - g_1) + \dots + \bar{d}_k g_k (1 - g_1)(1 - g_2) \dots (1 - g_{k-1})) \\ &\leq \frac{1}{\mathbb{P}[\Lambda^\nu]} \cdot \sum_{s=1}^k \bar{d}_s g_s \cdot \left(\sum_{t=1}^k g_t \prod_{z=1}^{t-1} (1 - g_z) \right) \end{aligned} \quad (5.3)$$

$$= \sum_{s=1}^k \bar{d}_s g_s \quad (5.4)$$

$$= C_\nu^{\text{avg}}. \quad (5.5)$$

The proof for inequality (5.3) is given in A.2 (note that $\sum_{s=1}^k g_s = 1$), equality (5.4) follows from $\mathbb{P}[\Lambda^\nu] = 1 - \prod_{t=1}^k (1 - g_t) = \sum_{t=1}^k g_t \prod_{z=1}^{t-1} (1 - g_z)$, and (5.5) follows from the definition of the distances \bar{d}_s , probabilities g_s , and simple algebra. ■

Next, we show an estimate on the probability that none of ν 's neighbors is opened by the algorithm.

Lemma 18 *The probability that none of ν 's neighbors is opened satisfies $\mathbb{P}[\neg\Lambda^\nu] \leq 1/e$.*

Proof. We use the same partition of $\overline{N}(\nu)$ into groups G_1, \dots, G_k as in the proof of Lemma 17. Denoting by g_s the probability that a group G_s is selected (and thus that it has an open facility), we have

$$\mathbb{P}[\neg\Lambda^\nu] = \prod_{s=1}^k (1 - g_s) \leq e^{-\sum_{s=1}^k g_s} = e^{-\sum_{\mu \in \overline{N}(\nu)} \bar{y}_\mu} = \frac{1}{e}.$$

In this derivation, we first use that $1 - x \leq e^{-x}$ holds for all x , the second equality follows from $\sum_{s=1}^k g_s = \sum_{\mu \in \overline{N}(\nu)} \bar{y}_\mu$ and the last equality follows from $\sum_{\mu \in \overline{N}(\nu)} \bar{y}_\mu = 1$. ■

We are now ready to estimate the unconditional expected connection cost of ν (in the case when $\overline{N}(\kappa) \setminus \overline{N}(\nu) \neq \emptyset$) as follows:

$$\begin{aligned} \mathbb{E}[C_\nu] &= \mathbb{E}[C_\nu \mid \Lambda^\nu] \cdot \mathbb{P}[\Lambda^\nu] + \mathbb{E}[C_\nu \mid \neg\Lambda^\nu] \cdot \mathbb{P}[\neg\Lambda^\nu] \\ &\leq C_\nu^{\text{avg}} \cdot \mathbb{P}[\Lambda^\nu] + (C_\nu^{\text{avg}} + 2\alpha_\nu^*) \cdot \mathbb{P}[\neg\Lambda^\nu] \end{aligned} \tag{5.6}$$

$$\begin{aligned} &= C_\nu^{\text{avg}} + 2\alpha_\nu^* \cdot \mathbb{P}[\neg\Lambda^\nu] \\ &\leq C_\nu^{\text{avg}} + \frac{2}{e} \cdot \alpha_\nu^*. \end{aligned} \tag{5.7}$$

In the above derivation, inequality (5.6) follows from Lemmas 16 and 17, and inequality (5.7) follows from Lemma 18.

We have thus shown that the bound (5.2) holds in all three cases. Summing over all demands ν of a client j , we can now bound the expected connection cost of client j :

$$\mathbb{E}[C_j] = \sum_{\nu \in j} \mathbb{E}[C_\nu] \leq \sum_{\nu \in j} (C_\nu^{\text{avg}} + \frac{2}{e} \cdot \alpha_\nu^*) = C_j^* + \frac{2}{e} \cdot r_j \alpha_j^*.$$

Finally, summing over all clients j , we obtain our bound on the expected connection cost,

$$\mathbb{E}[C_{\text{ECHS}}] \leq C^* + \frac{2}{e} \cdot \text{LP}^*.$$

Therefore we have established that our algorithm constructs a feasible integral solution with an overall expected cost

$$\mathbb{E}[F_{\text{ECHS}} + C_{\text{ECHS}}] \leq F^* + C^* + \frac{2}{e} \cdot \text{LP}^* = (1 + 2/e) \cdot \text{LP}^* \leq (1 + 2/e) \cdot \text{OPT}.$$

Summarizing, we obtain the main result of this section.

Theorem 19 *Algorithm ECHS is a $(1 + 2/e)$ -approximation algorithm for FTFP.*

5.3 Algorithm EBGs with Ratio 1.575

In this section we give our main result, a 1.575-approximation algorithm for FTFP, where 1.575 is the value of $\min_{\gamma \geq 1} \max\{\gamma, 1 + 2/e^\gamma, \frac{1/e+1/e^\gamma}{1-1/\gamma}\}$, rounded to three decimal digits. This matches the ratio of the best known LP-rounding algorithm for UFL by Byrka *et al.* [5].

Recall that in Section 5.2 we showed how to compute an integral solution with facility cost bounded by F^* and connection cost bounded by $C^* + 2/e \cdot \text{LP}^*$. Thus, while our facility cost does not exceed the optimal fractional facility cost, our connection cost is significantly larger than the connection cost in the optimal fractional solution. A natural idea is to balance these two ratios by reducing the connection cost at the expense of the facility cost. One way to do this would be to increase the probability of opening facilities, from \bar{y}_μ (used in Algorithm ECHS) to, say, $\gamma \bar{y}_\mu$, for some $\gamma > 1$. This increases the expected

facility cost by a factor of γ but, as it turns out, it also reduces the probability that an indirect connection occurs for a non-primary demand to $1/e^\gamma$ (from the previous value $1/e$ in ECHS). As a consequence, for each primary demand κ , the new algorithm will select a facility to open from the nearest facilities μ in $\overline{N}(\kappa)$ such that the connection values $\bar{x}_{\mu\nu}$ sum up to $1/\gamma$, instead of 1 as in Algorithm ECHS. It is easily seen that this will improve the estimate on connection cost for primary demands. These two changes, along with a more refined analysis, are the essence of the approach in [5], expressed in our terminology.

Our approach can be thought of as a combination of the above ideas with the techniques of demand reduction and adaptive partitioning that we introduced earlier. However, our adaptive partitioning technique needs to be carefully modified, because now we will be using a more intricate neighborhood structure, with the neighborhood of each demand divided into two disjoint parts, and with restrictions on how parts from different demands can overlap.

We begin by describing properties that our partitioned fractional solution (\bar{x}, \bar{y}) needs to satisfy. Assume that γ is some constant such that $1 < \gamma < 2$. As mentioned earlier, the neighborhood $\overline{N}(\nu)$ of each demand ν will be divided into two disjoint parts. The first part, called the *close neighborhood* and denoted $\overline{N}_{\text{cls}}(\nu)$, contains the facilities in $\overline{N}(\nu)$ nearest to ν with the total connection value equal $1/\gamma$, that is $\sum_{\mu \in \overline{N}_{\text{cls}}(\nu)} \bar{x}_{\mu\nu} = 1/\gamma$. The second part, called the *far neighborhood* and denoted $\overline{N}_{\text{far}}(\nu)$, contains the remaining facilities in $\overline{N}(\nu)$ (so $\sum_{\mu \in \overline{N}_{\text{far}}(\nu)} \bar{x}_{\mu\nu} = 1 - 1/\gamma$). We restate these definitions formally below in Property (NB). Recall that for any set A of facilities and a demand ν , by $D(A, \nu)$ we denote the average distance between ν and the facilities in A , that is $D(A, \nu) = \sum_{\mu \in A} d_{\mu\nu} \bar{y}_\mu / \sum_{\mu \in A} \bar{y}_\mu$.

We will use notations $C_{\text{cls}}^{\text{avg}}(\nu) = D(\overline{N}_{\text{cls}}(\nu), \nu)$ and $C_{\text{far}}^{\text{avg}}(\nu) = D(\overline{N}_{\text{far}}(\nu), \nu)$ for the average distances from ν to its close and far neighborhoods, respectively. By the definition of these sets and the completeness property (CO), these distances can be expressed as

$$C_{\text{cls}}^{\text{avg}}(\nu) = \gamma \sum_{\mu \in \overline{N}_{\text{cls}}(\nu)} d_{\mu\nu} \bar{x}_{\mu\nu} \quad \text{and} \quad C_{\text{far}}^{\text{avg}}(\nu) = \frac{\gamma}{\gamma - 1} \sum_{\mu \in \overline{N}_{\text{far}}(\nu)} d_{\mu\nu} \bar{x}_{\mu\nu}.$$

We will also use notation $C_{\text{cls}}^{\text{max}}(\nu) = \max_{\mu \in \overline{N}_{\text{cls}}(\nu)} d_{\mu\nu}$ for the maximum distance from ν to its close neighborhood. The average distance from a demand ν to its overall neighborhood $\overline{N}(\nu)$ is denoted as $C^{\text{avg}}(\nu) = D(\overline{N}(\nu), \nu) = \sum_{\mu \in \overline{N}(\nu)} d_{\mu\nu} \bar{x}_{\mu\nu}$. It is easy to see that

$$C^{\text{avg}}(\nu) = \frac{1}{\gamma} C_{\text{cls}}^{\text{avg}}(\nu) + \frac{\gamma - 1}{\gamma} C_{\text{far}}^{\text{avg}}(\nu). \quad (5.8)$$

Our partitioned solution $(\bar{\mathbf{x}}, \bar{\mathbf{y}})$ must satisfy the same partitioning and completeness properties as before, namely properties (PS) and (CO) in Section 4.2. In addition, it must satisfy a new neighborhood property (NB) and modified properties (PD') and (SI'), listed below.

(NB) *Neighborhoods*. For each demand $\nu \in \overline{\mathcal{C}}$, its neighborhood is divided into *close* and *far* neighborhood, that is $\overline{N}(\nu) = \overline{N}_{\text{cls}}(\nu) \cup \overline{N}_{\text{far}}(\nu)$, where

- $\overline{N}_{\text{cls}}(\nu) \cap \overline{N}_{\text{far}}(\nu) = \emptyset$,
- $\sum_{\mu \in \overline{N}_{\text{cls}}(\nu)} \bar{x}_{\mu\nu} = 1/\gamma$, and
- if $\mu \in \overline{N}_{\text{cls}}(\nu)$ and $\mu' \in \overline{N}_{\text{far}}(\nu)$ then $d_{\mu\nu} \leq d_{\mu'\nu}$.

Note that the first two conditions, together with (PS.1), imply that $\sum_{\mu \in \overline{N}_{\text{far}}(\nu)} \bar{x}_{\mu\nu} = 1 - 1/\gamma$. When defining $\overline{N}_{\text{cls}}(\nu)$, in case of ties, which can occur when some facilities

in $\overline{N}(\nu)$ are at the same distance from ν , we use a tie-breaking rule that is explained in the proof of Lemma 20 (the only place where the rule is needed).

(PD') *Primary demands.* Primary demands satisfy the following conditions:

1. For any two different primary demands $\kappa, \kappa' \in P$ we have $\overline{N}_{\text{cls}}(\kappa) \cap \overline{N}_{\text{cls}}(\kappa') = \emptyset$.
2. For each site $i \in \mathbb{F}$, $\sum_{\kappa \in P} \sum_{\mu \in i \cap \overline{N}_{\text{cls}}(\kappa)} \bar{x}_{\mu\kappa} \leq y_i^*$. In the summation, as before, we overload notation i to stand for the set of facilities created on site i .
3. Each demand $\nu \in \overline{\mathbb{C}}$ is assigned to one primary demand $\kappa \in P$ such that
 - (a) $\overline{N}_{\text{cls}}(\nu) \cap \overline{N}_{\text{cls}}(\kappa) \neq \emptyset$, and
 - (b) $C_{\text{cls}}^{\text{avg}}(\nu) + C_{\text{cls}}^{\text{max}}(\nu) \geq C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{cls}}^{\text{max}}(\kappa)$.

(SI') *Siblings.* For any pair $\nu, \nu' \in \overline{\mathbb{C}}$ of different siblings we have

1. $\overline{N}(\nu) \cap \overline{N}(\nu') = \emptyset$.
2. If ν is assigned to a primary demand κ then $\overline{N}(\nu') \cap \overline{N}_{\text{cls}}(\kappa) = \emptyset$. In particular, by Property (PD'.3(a)), this implies that different sibling demands are assigned to different primary demands, since $\overline{N}_{\text{cls}}(\nu')$ is a subset of $\overline{N}(\nu')$.

Modified adaptive partitioning. To obtain a fractional solution with the above properties, we employ a modified adaptive partitioning algorithm. As in Section 4.2, we have two phases. In Phase 1 we split clients into demands and create facilities on sites, while in Phase 2 we augment each demand's connection values $\bar{x}_{\mu\nu}$ so that the total connection value of each demand ν is 1. As the partitioning algorithm proceeds, for any demand ν , $\overline{N}(\nu)$ denotes the set of facilities with $\bar{x}_{\mu\nu} > 0$; hence the notation $\overline{N}(\nu)$ actually represents

a dynamic set which gets fixed once the partitioning algorithm concludes both Phase 2. On the other hand, $\overline{N}_{\text{cls}}(\nu)$ and $\overline{N}_{\text{far}}(\nu)$ refer to the close and far neighborhoods at the time when $\overline{N}(\nu)$ is fixed.

Similar to the algorithm in Section 4.2, Phase 1 runs in iterations. Fix some iteration and consider any client j . As before, $\tilde{N}(j)$ is the neighborhood of j with respect to the yet unpartitioned solution, namely the set of facilities μ such that $\tilde{x}_{\mu j} > 0$. Order the facilities in this set as $\tilde{N}(j) = \{\mu_1, \dots, \mu_q\}$ with non-decreasing distance from j , that is $d_{\mu_1 j} \leq d_{\mu_2 j} \leq \dots \leq d_{\mu_q j}$. Without loss of generality, there is an index l for which $\sum_{s=1}^l \tilde{x}_{\mu_s j} = 1/\gamma$, since we can always split one facility to achieve this. Then we define $\tilde{N}_{\text{cls}}(j) = \{\mu_1, \dots, \mu_l\}$. (Unlike close neighborhoods of demands, $\tilde{N}_{\text{cls}}(j)$ can vary over time.)

We also use notation

$$\text{tcc}_{\text{cls}}(j) = D(\tilde{N}_{\text{cls}}(j), j) = \gamma \sum_{\mu \in \tilde{N}_{\text{cls}}(j)} d_{\mu j} \tilde{x}_{\mu j} \quad \text{and} \quad \text{dmax}_{\text{cls}}(j) = \max_{\mu \in \tilde{N}_{\text{cls}}(j)} d_{\mu j}.$$

When the iteration starts, we first find a not-yet-exhausted client p that minimizes the value of $\text{tcc}_{\text{cls}}(p) + \text{dmax}_{\text{cls}}(p)$ and create a new demand ν for p . Now we have two cases:

Case 1: $\tilde{N}_{\text{cls}}(p) \cap \overline{N}(\kappa) \neq \emptyset$ for some existing primary demand $\kappa \in P$. In this case we assign ν to κ . As before, if there are multiple such κ , we pick any of them. We also fix $\bar{x}_{\mu\nu} \leftarrow \tilde{x}_{\mu p}$ and $\tilde{x}_{\mu p} \leftarrow 0$ for each $\mu \in \tilde{N}(p) \cap \overline{N}(\kappa)$. Note that although we check for overlap between $\tilde{N}_{\text{cls}}(p)$ and $\overline{N}(\kappa)$, the facilities we actually move into $\overline{N}(\nu)$ include all facilities in the intersection of $\tilde{N}(p)$, a bigger set, with $\overline{N}(\kappa)$.

At this time, the total connection value between ν and $\mu \in \overline{N}(\nu)$ is at most $1/\gamma$, since

$\sum_{\mu \in \overline{N}(\nu)} \bar{y}_{\mu} = 1/\gamma$ (this follows from the definition of neighborhoods for new primary

demands in Case 2 below) and we have $\overline{N}(\nu) \subseteq \overline{N}(\kappa)$ at this point. Later in Phase 2 we will add additional facilities from $\tilde{N}(p)$ to $\overline{N}(\nu)$ to make ν 's total connection value equal to 1.

Case 2: $\tilde{N}_{\text{cls}}(p) \cap \overline{N}(\kappa) = \emptyset$ for all existing primary demands $\kappa \in P$. In this case we make

ν a primary demand (that is, add it to P) and assign it to itself. We then move the facilities from $\tilde{N}_{\text{cls}}(p)$ to $\overline{N}(\nu)$, that is for $\mu \in \tilde{N}_{\text{cls}}(p)$ we set $\bar{x}_{\mu\nu} \leftarrow \tilde{x}_{\mu p}$ and $\tilde{x}_{\mu p} \leftarrow 0$.

It is easy to see that the total connection value of ν to $\overline{N}(\nu)$ is now exactly $1/\gamma$, that is $\sum_{\mu \in \overline{N}(\nu)} \bar{y}_\mu = 1/\gamma$. Moreover, facilities remaining in $\tilde{N}(p)$ are all farther away from ν than those in $\overline{N}(\nu)$. As we add only facilities from $\tilde{N}(p)$ to $\overline{N}(\nu)$ in Phase 2, the final $\overline{N}_{\text{cls}}(\nu)$ contains the same set of facilities as the current set $\overline{N}(\nu)$. (More precisely, $\overline{N}_{\text{cls}}(\nu)$ consists of the facilities that either are currently in $\overline{N}(\nu)$ or were obtained from splitting the facilities currently in $\overline{N}(\nu)$.)

Once all clients are exhausted, that is, each client j has r_j demands created, Phase 1 concludes. We then run Phase 2, the augmenting phase, following the same steps as in Section 4.2. For each client j and each demand $\nu \in j$ with total connection value to $\overline{N}(\nu)$ less than 1 (that is, $\sum_{\mu \in \overline{N}(\nu)} \bar{x}_{\mu\nu} < 1$), we use our `AUGMENTTOUNIT()` procedure to add additional facilities (possibly split, if necessary) from $\tilde{N}(j)$ to $\overline{N}(\nu)$ to make the total connection value between ν and $\overline{N}(\nu)$ equal 1.

This completes the description of the partitioning algorithm. Summarizing, for each client $j \in \mathbb{C}$ we created r_j demands on the same point as j , and we created a number of facilities at each site $i \in \mathbb{F}$. Thus computed sets of demands and facilities are denoted

$\overline{\mathbb{C}}$ and $\overline{\mathbb{F}}$, respectively. For each facility $\mu \in i$ we defined its fractional opening value \bar{y}_μ , $0 \leq \bar{y}_\mu \leq 1$, and for each demand $\nu \in j$ we defined its fractional connection value $\bar{x}_{\mu\nu} \in \{0, \bar{y}_\mu\}$. The connections with $\bar{x}_{\mu\nu} > 0$ define the neighborhood $\overline{N}(\nu)$. The facilities in $\overline{N}(\nu)$ that are closest to ν and have total connection value from ν equal $1/\gamma$ form the close neighborhood $\overline{N}_{\text{cls}}(\nu)$, while the remaining facilities in $\overline{N}(\nu)$ form the far neighborhood $\overline{N}_{\text{far}}(\nu)$. It remains to show that this partitioning satisfies all the desired properties.

Correctness of partitioning. We now argue that our partitioned fractional solution (\bar{x}, \bar{y}) satisfies all the stated properties. Properties (PS), (CO) and (NB) are directly enforced by the algorithm.

(PD'.1) holds because for each primary demand $\kappa \in p$, $\overline{N}_{\text{cls}}(\kappa)$ is the same set as $\tilde{N}_{\text{cls}}(p)$ at the time when κ was created, and $\tilde{N}_{\text{cls}}(p)$ is removed from $\tilde{N}(p)$ right after this step. Further, the partitioning algorithm makes κ a primary demand only if $\tilde{N}_{\text{cls}}(p)$ is disjoint from the set $\overline{N}(\kappa')$ of all existing primary demands κ' at that iteration, but these neighborhoods are the same as the final close neighborhoods $\overline{N}_{\text{cls}}(\kappa')$.

The justification of (PD'.2) is similar to that for (PD.2) from Section 4.2. All close neighborhoods of primary demands are disjoint, due to (PD'.1), so each facility $\mu \in i$ can appear in at most one $\overline{N}_{\text{cls}}(\kappa)$, for some $\kappa \in P$. Condition (CO) implies that $\bar{y}_\mu = \bar{x}_{\mu\kappa}$ for $\mu \in \overline{N}_{\text{cls}}(\kappa)$. As a result, the summation on the left-hand side is not larger than $\sum_{\mu \in i} \bar{y}_\mu = y_i^*$.

Regarding (PD'.3(a)), at first glance this property seems to follow directly from the algorithm, as we only assign a demand ν to a primary demand κ when $\overline{N}(\nu)$ at that iteration overlaps with $\overline{N}(\kappa)$ (which is equal to the final value of $\overline{N}_{\text{cls}}(\kappa)$). However, it is

a little more subtle, as the final $\overline{N}_{\text{cls}}(\nu)$ may contain facilities added to $\overline{N}(\nu)$ in Phase 2. Those facilities may turn out to be closer to ν than some facilities in $\overline{N}(\kappa) \cap \tilde{N}(j)$ (not $\tilde{N}_{\text{cls}}(j)$) that we added to $\overline{N}(\nu)$ in Phase 1. If the final $\overline{N}_{\text{cls}}(\nu)$ consists only of facilities added in Phase 2, we no longer have the desired overlap of $\overline{N}_{\text{cls}}(\kappa)$ and $\overline{N}_{\text{cls}}(\nu)$. Luckily this bad scenario never occurs. We postpone the proof of this property to Lemma 20. The proof of (PD'.3(b)) is similar to that of Lemma 12, and we defer it to Lemma 21.

(SI'.1) follows directly from the algorithm because for each demand $\nu \in j$, all facilities added to $\overline{N}(\nu)$ are immediately removed from $\tilde{N}(j)$ and each facility is added to $\overline{N}(\nu)$ of exactly one demand $\nu \in j$. Splitting facilities obviously preserves (SI'.1).

The proof of (SI'.2) is similar to that of Lemma 10. If $\kappa = \nu$ then (SI'.2) follows from (SI'.1), so we can assume that $\kappa \neq \nu$. Suppose that $\nu' \in j$ is assigned to $\kappa' \in P$ and consider the situation after Phase 1. By the way we reassign facilities in Case 1, at this time we have $\overline{N}(\nu) \subseteq \overline{N}(\kappa) = \overline{N}_{\text{cls}}(\kappa)$ and $\overline{N}(\nu') \subseteq \overline{N}(\kappa') = \overline{N}_{\text{cls}}(\kappa')$, so $\overline{N}(\nu') \cap \overline{N}_{\text{cls}}(\kappa) = \emptyset$, by (PD'.1). Moreover, we have $\tilde{N}(j) \cap \overline{N}_{\text{cls}}(\kappa) = \emptyset$ after this iteration, because any facilities that were also in $\overline{N}_{\text{cls}}(\kappa)$ were removed from $\tilde{N}(j)$ when ν was created. In Phase 2, augmentation does not change $\overline{N}_{\text{cls}}(\kappa)$ and all facilities added to $\overline{N}(\nu')$ are from the set $\tilde{N}(j)$ at the end of Phase 1, which is a subset of the set $\tilde{N}(j)$ after this iteration, since $\tilde{N}(j)$ can only shrink. So the condition (SI'.2) will remain true.

Lemma 20 *Property (PD'.3(a)) holds.*

Proof. Let j be the client for which $\nu \in j$. We consider an iteration when we create ν from j and assign it to κ , and within this proof, notation $\tilde{N}_{\text{cls}}(j)$ and $\tilde{N}(j)$ will refer to the value of the sets at this particular time. At this time, $\overline{N}(\nu)$ is initialized to

$\tilde{N}(j) \cap \overline{N}(\kappa)$. Recall that $\overline{N}(\kappa)$ is now equal to the final $\overline{N}_{\text{cls}}(\kappa)$ (taking into account facility splitting). We would like to show that the set $\tilde{N}_{\text{cls}}(j) \cap \overline{N}_{\text{cls}}(\kappa)$ (which is not empty) will be included in $\overline{N}_{\text{cls}}(\nu)$ at the end. Technically speaking, this will not be true due to facility splitting, so we need to rephrase this claim and the proof in terms of the set of facilities obtained after the algorithm completes.

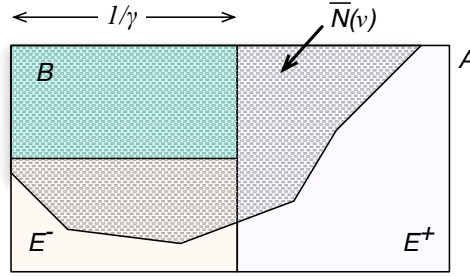


Figure 5.1: Illustration of the sets $\overline{N}(\nu)$, A , B , E^- and E^+ in the proof of Lemma 20. Let $X \subseteq Y$ mean that the facility sets X is obtained from Y by splitting facilities. We then have $A \subseteq \tilde{N}(j)$, $B \subseteq \tilde{N}_{\text{cls}}(j) \cap \overline{N}_{\text{cls}}(\kappa)$, $E^- \subseteq \tilde{N}_{\text{cls}}(j) - \overline{N}_{\text{cls}}(\kappa)$, $E^+ \subseteq \tilde{N}(j) - \tilde{N}_{\text{cls}}(j)$.

We define the sets A , B , E^- and E^+ as the subsets of $\overline{\mathbb{F}}$ (the final set of facilities) that were obtained from splitting facilities in the sets $\tilde{N}(j)$, $\tilde{N}_{\text{cls}}(j) \cap \overline{N}_{\text{cls}}(\kappa)$, $\tilde{N}_{\text{cls}}(j) - \overline{N}_{\text{cls}}(\kappa)$ and $\tilde{N}(j) - \tilde{N}_{\text{cls}}(j)$, respectively. (See Figure 5.1.) We claim that at the end $B \subseteq \overline{N}_{\text{cls}}(\nu)$, with the caveat that the ties in the definition of $\overline{N}_{\text{cls}}(\nu)$ are broken in favor of the facilities in B . (This is the tie-breaking rule that we mentioned in the definition of $\overline{N}_{\text{cls}}(\nu)$.) This will be sufficient to prove the lemma because $B \neq \emptyset$, by the algorithm.

We now prove this claim. In this paragraph $\overline{N}(\nu)$ denotes the final set $\overline{N}(\nu)$ after both phases are completed. Thus the total connection value of $\overline{N}(\nu)$ to ν is 1. Note first that

$B \subseteq \overline{N}(\nu) \subseteq A$, because we never remove facilities from $\overline{N}(\nu)$ and we only add facilities from $\tilde{N}(j)$. Also, $B \cup E^-$ represents the facilities obtained from $\tilde{N}_{\text{cls}}(j)$, so $\sum_{\mu \in B \cup E^-} \bar{y}_\mu = 1/\gamma$. This and $B \subseteq \overline{N}(\nu)$ implies that the total connection value of $B \cup (\overline{N}(\nu) \cap E^-)$ to ν is at most $1/\gamma$. But all facilities in $B \cup (\overline{N}(\nu) \cap E^-)$ are closer to ν (taking into account our tie breaking in property (NB)) than those in $E^+ \cap \overline{N}(\nu)$. It follows that $B \subseteq \overline{N}_{\text{cls}}(\nu)$, completing the proof. ■

Lemma 21 *Property (PD'.3(b)) holds.*

Proof. This proof is similar to that for Lemma 12. For a client j and demand η , we will write $\text{tcc}_{\text{cls}}^\eta(j)$ and $\text{dmax}_{\text{cls}}^\eta(j)$ to denote the values of $\text{tcc}_{\text{cls}}(j)$ and $\text{dmax}_{\text{cls}}(j)$ at the time when η was created. (Here η may or may not be a demand of client j).

Suppose $\nu \in j$ is assigned to a primary demand $\kappa \in p$. By the way primary demands are constructed in the partitioning algorithm, $\tilde{N}_{\text{cls}}(p)$ becomes $\overline{N}(\kappa)$, which is equal to the final value of $\overline{N}_{\text{cls}}(\kappa)$. So we have $C_{\text{cls}}^{\text{avg}}(\kappa) = \text{tcc}_{\text{cls}}^\kappa(p)$ and $C_{\text{cls}}^{\text{max}}(\kappa) = \text{dmax}_{\text{cls}}^\kappa(p)$. Further, since we choose p to minimize $\text{tcc}_{\text{cls}}(p) + \text{dmax}_{\text{cls}}(p)$, we have that $\text{tcc}_{\text{cls}}^\kappa(p) + \text{dmax}_{\text{cls}}^\kappa(p) \leq \text{tcc}_{\text{cls}}^\kappa(j) + \text{dmax}_{\text{cls}}^\kappa(j)$.

Using an argument analogous to that in the proof of Lemma 11, our modified partitioning algorithm guarantees that $\text{tcc}_{\text{cls}}^\kappa(j) \leq \text{tcc}_{\text{cls}}^\nu(j) \leq C_{\text{cls}}^{\text{avg}}(\nu)$ and $\text{dmax}_{\text{cls}}^\kappa(j) \leq \text{dmax}_{\text{cls}}^\nu(j) \leq C_{\text{cls}}^{\text{max}}(\nu)$ since ν was created later. Therefore, we have

$$\begin{aligned} C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{cls}}^{\text{max}}(\kappa) &= \text{tcc}_{\text{cls}}^\kappa(p) + \text{dmax}_{\text{cls}}^\kappa(p) \\ &\leq \text{tcc}_{\text{cls}}^\kappa(j) + \text{dmax}_{\text{cls}}^\kappa(j) \leq \text{tcc}_{\text{cls}}^\nu(j) + \text{dmax}_{\text{cls}}^\nu(j) \leq C_{\text{cls}}^{\text{avg}}(\nu) + C_{\text{cls}}^{\text{max}}(\nu), \end{aligned}$$

completing the proof. ■

Now we have completed the proof that the computed partitioning satisfies all the required properties.

Algorithm EBGs. The complete algorithm starts with solving the LP(3.1) and computing the partitioning described earlier in this section. Given the partitioned fractional solution (\bar{x}, \bar{y}) with the desired properties, we start the process of opening facilities and making connections to obtain an integral solution. To this end, for each primary demand $\kappa \in P$, we open exactly one facility $\phi(\kappa)$ in $\bar{N}_{\text{cls}}(\kappa)$, where each $\mu \in \bar{N}_{\text{cls}}(\kappa)$ is chosen as $\phi(\kappa)$ with probability $\gamma \bar{y}_\mu$. For all facilities $\mu \in \bar{\mathbb{F}} - \bigcup_{\kappa \in P} \bar{N}_{\text{cls}}(\kappa)$, we open them independently, each with probability $\gamma \bar{y}_\mu$.

We claim that all probabilities are well-defined, that is $\gamma \bar{y}_\mu \leq 1$ for all μ . Indeed, if $\bar{y}_\mu > 0$ then $\bar{y}_\mu = \bar{x}_{\mu\nu}$ for some ν , by Property (CO). If $\mu \in \bar{N}_{\text{cls}}(\nu)$ then the definition of close neighborhoods implies that $\bar{x}_{\mu\nu} \leq 1/\gamma$. If $\mu \in \bar{N}_{\text{far}}(\nu)$ then $\bar{x}_{\mu\nu} \leq 1 - 1/\gamma \leq 1/\gamma$, because $\gamma < 2$. Thus $\gamma \bar{y}_\mu \leq 1$, as claimed.

Next, we connect demands to facilities. Each primary demand $\kappa \in P$ will connect to the only open facility $\phi(\kappa)$ in $\bar{N}_{\text{cls}}(\kappa)$. For each non-primary demand $\nu \in \bar{\mathbb{C}} - P$, if there is an open facility in $\bar{N}_{\text{cls}}(\nu)$ then we connect ν to the nearest such facility. Otherwise, we connect ν to the nearest far facility in $\bar{N}_{\text{far}}(\nu)$ if one is open. Otherwise, we connect ν to its *target facility* $\phi(\kappa)$, where κ is the primary demand that ν is assigned to.

Analysis. By the algorithm, for each client j , all its r_j demands are connected to open facilities. If two different siblings $\nu, \nu' \in j$ are assigned, respectively, to primary demands κ, κ' then, by Properties (SI'.1), (SI'.2), and (PD'.1) we have

$$(\bar{N}(\nu) \cup \bar{N}_{\text{cls}}(\kappa)) \cap (\bar{N}(\nu') \cup \bar{N}_{\text{cls}}(\kappa')) = \emptyset.$$

This condition guarantees that ν and ν' are assigned to different facilities, regardless whether they are connected to a neighbor facility or to its target facility. Therefore the computed solution is feasible.

We now estimate the cost of the solution computed by Algorithm EBGs. The lemma below bounds the expected facility cost.

Lemma 22 *The expectation of facility cost F_{EBGS} of Algorithm EBGs is at most γF^* .*

Proof. By the algorithm, each facility $\mu \in \bar{\mathbb{F}}$ is opened with probability $\gamma \bar{y}_\mu$, independently of whether it belongs to the close neighborhood of a primary demand or not. Therefore, by linearity of expectation, we have that the expected facility cost is

$$\mathbb{E}[F_{\text{EBGS}}] = \sum_{\mu \in \bar{\mathbb{F}}} f_\mu \gamma \bar{y}_\mu = \gamma \sum_{i \in \mathbb{F}} f_i \sum_{\mu \in i} \bar{y}_\mu = \gamma \sum_{i \in \mathbb{F}} f_i y_i^* = \gamma F^*,$$

where the third equality follows from (PS.3). ■

In the remainder of this section we focus on the connection cost. Let C_ν be the random variable representing the connection cost of a demand ν . Our objective is to show that the expectation of ν satisfies

$$\mathbb{E}[C_\nu] \leq C^{\text{avg}}(\nu) \cdot \max \left\{ \frac{1/e + 1/e^\gamma}{1 - 1/\gamma}, 1 + \frac{2}{e^\gamma} \right\}. \quad (5.9)$$

If ν is a primary demand then, due to the algorithm, we have $\mathbb{E}[C_\nu] = C_{\text{cls}}^{\text{avg}}(\nu) \leq C^{\text{avg}}(\nu)$, so (5.9) is easily satisfied.

Thus for the rest of the argument we will focus on the case when ν is a non-primary demand. Recall that the algorithm connects ν to the nearest open facility in $\bar{N}_{\text{cls}}(\nu)$ if at least one facility in $\bar{N}_{\text{cls}}(\nu)$ is open. Otherwise the algorithm connects ν to the nearest open

facility in $\overline{N}_{\text{far}}(\nu)$, if any. In the event that no facility in $\overline{N}(\nu)$ opens, the algorithm will connect ν to its target facility $\phi(\kappa)$, where κ is the primary demand that ν was assigned to, and $\phi(\kappa)$ is the only facility open in $\overline{N}_{\text{cls}}(\kappa)$. Let Λ^ν denote the event that at least one facility in $\overline{N}(\nu)$ is open and Λ_{cls}^ν be the event that at least one facility in $\overline{N}_{\text{cls}}(\nu)$ is open. $\neg\Lambda^\nu$ denotes the complement event of Λ^ν , that is, the event that none of ν 's neighbors opens. We want to estimate the following three conditional expectations:

$$\mathbb{E}[C_\nu \mid \Lambda_{\text{cls}}^\nu], \quad \mathbb{E}[C_\nu \mid \Lambda^\nu \wedge \neg\Lambda_{\text{cls}}^\nu], \quad \text{and} \quad \mathbb{E}[C_\nu \mid \neg\Lambda^\nu],$$

and their associated probabilities.

We start with a lemma dealing with the third expectation, $\mathbb{E}[C_\nu \mid \neg\Lambda^\nu] = \mathbb{E}[d_{\phi(\kappa)\nu} \mid \Lambda^\nu]$. The proof of this lemma relies on Properties (PD'.3(a)) and (PD'.3(b)) of modified partitioning and follows the reasoning in the proof of a similar lemma in [5, 4].

Lemma 23 *Assuming that no facility in $\overline{N}(\nu)$ opens, the expected connection cost of ν is*

$$\mathbb{E}[C_\nu \mid \neg\Lambda^\nu] \leq C_{\text{cls}}^{\text{avg}}(\nu) + 2C_{\text{far}}^{\text{avg}}(\nu). \quad (5.10)$$

Proof. It suffices to show a stronger inequality

$$\mathbb{E}[C_\nu \mid \neg\Lambda^\nu] \leq C_{\text{cls}}^{\text{avg}}(\nu) + C_{\text{cls}}^{\text{max}}(\nu) + C_{\text{far}}^{\text{avg}}(\nu), \quad (5.11)$$

which then implies (5.10) because $C_{\text{cls}}^{\text{max}}(\nu) \leq C_{\text{far}}^{\text{avg}}(\nu)$. The proof of (5.11) is similar to that in [4]. For the sake of completeness, we provide it here, formulated in our terminology and notation.

Assume that the event $\neg\Lambda^\nu$ is true, that is Algorithm EBGs does not open any facility in $\overline{N}(\nu)$. Let κ be the primary demand that ν was assigned to. Also let

$$K = \overline{N}_{\text{cls}}(\kappa) \setminus \overline{N}(\nu), \quad V_{\text{cls}} = \overline{N}_{\text{cls}}(\kappa) \cap \overline{N}_{\text{cls}}(\nu) \quad \text{and} \quad V_{\text{far}} = \overline{N}_{\text{cls}}(\kappa) \cap \overline{N}_{\text{far}}(\nu).$$

Then $K, V_{\text{cls}}, V_{\text{far}}$ form a partition of $\overline{N}_{\text{cls}}(\kappa)$, that is, they are disjoint and their union is $\overline{N}_{\text{cls}}(\kappa)$. Moreover, we have that K is not empty, because Algorithm EBGs opens some facility in $\overline{N}_{\text{cls}}(\kappa)$ and this facility cannot be in $V_{\text{cls}} \cup V_{\text{far}}$, by our assumption. We also have that V_{cls} is not empty due to (PD'.3(a)).

Recall that $D(A, \eta) = \sum_{\mu \in A} d_{\mu\eta} \bar{y}_\mu / \sum_{\mu \in A} \bar{y}_\mu$ is the average distance between a demand η and the facilities in a set A . We shall show that

$$D(K, \nu) \leq C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu). \quad (5.12)$$

This is sufficient, because, by the algorithm, $D(K, \nu)$ is exactly the expected connection cost for demand ν conditioned on the event that none of ν 's neighbors opens, that is the left-hand side of (5.11). Further, (PD'.3(b)) states that $C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{cls}}^{\text{max}}(\kappa) \leq C_{\text{cls}}^{\text{avg}}(\nu) + C_{\text{cls}}^{\text{max}}(\nu)$, and thus (5.12) implies (5.11).

The proof of (5.12) is by analysis of several cases.

Case 1: $D(K, \kappa) \leq C_{\text{cls}}^{\text{avg}}(\kappa)$. For any facility $\mu \in V_{\text{cls}}$ (recall that $V_{\text{cls}} \neq \emptyset$), we have $d_{\mu\kappa} \leq C_{\text{cls}}^{\text{max}}(\kappa)$ and $d_{\mu\nu} \leq C_{\text{cls}}^{\text{max}}(\nu) \leq C_{\text{far}}^{\text{avg}}(\nu)$. Therefore, using the case assumption, we get $D(K, \nu) \leq D(K, \kappa) + d_{\mu\kappa} + d_{\mu\nu} \leq C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu)$.

Case 2: There exists a facility $\mu \in V_{\text{cls}}$ such that $d_{\mu\kappa} \leq C_{\text{cls}}^{\text{avg}}(\kappa)$. Since $\mu \in V_{\text{cls}}$, we infer that $d_{\mu\nu} \leq C_{\text{cls}}^{\text{max}}(\nu) \leq C_{\text{far}}^{\text{avg}}(\nu)$. Using $C_{\text{cls}}^{\text{max}}(\kappa)$ to bound $D(K, \kappa)$, we have $D(K, \nu) \leq D(K, \kappa) + d_{\mu\kappa} + d_{\mu\nu} \leq C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu)$.

Case 3: In this case we assume that neither of Cases 1 and 2 applies, that is $D(K, \kappa) > C_{\text{cls}}^{\text{avg}}(\kappa)$ and every $\mu \in V_{\text{cls}}$ satisfies $d_{\mu\kappa} > C_{\text{cls}}^{\text{avg}}(\kappa)$. This implies that $D(K \cup V_{\text{cls}}, \kappa) > C_{\text{cls}}^{\text{avg}}(\kappa) = D(\overline{N}_{\text{cls}}(\kappa), \kappa)$. Since sets K, V_{cls} and V_{far} form a partition of $\overline{N}_{\text{cls}}(\kappa)$, we obtain

that in this case V_{far} is not empty and $D(V_{\text{far}}, \kappa) < C_{\text{cls}}^{\text{avg}}(\kappa)$. Let $\delta = C_{\text{cls}}^{\text{avg}}(\kappa) - D(V_{\text{far}}, \kappa) > 0$.

We now have two sub-cases:

Case 3.1: $D(V_{\text{far}}, \nu) \leq C_{\text{far}}^{\text{avg}}(\nu) + \delta$. Substituting δ , this implies that $D(V_{\text{far}}, \nu) + D(V_{\text{far}}, \kappa) \leq$

$C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu)$. From the definition of the average distance $D(V_{\text{far}}, \kappa)$ and $D(V_{\text{far}}, \nu)$,

we obtain that there exists some $\mu \in V_{\text{far}}$ such that $d_{\mu\kappa} + d_{\mu\nu} \leq C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu)$.

Thus $D(K, \nu) \leq D(K, \kappa) + d_{\mu\kappa} + d_{\mu\nu} \leq C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu)$.

Case 3.2: $D(V_{\text{far}}, \nu) > C_{\text{far}}^{\text{avg}}(\nu) + \delta$. The case assumption implies that V_{far} is a proper

subset of $\bar{N}_{\text{far}}(\nu)$, that is $\bar{N}_{\text{far}}(\nu) \setminus V_{\text{far}} \neq \emptyset$. Let $\hat{y} = \gamma \sum_{\mu \in V_{\text{far}}} \bar{y}_{\mu}$. We can express

$C_{\text{far}}^{\text{avg}}(\nu)$ using \hat{y} as follows

$$C_{\text{far}}^{\text{avg}}(\nu) = D(V_{\text{far}}, \nu) \frac{\hat{y}}{\gamma - 1} + D(\bar{N}_{\text{far}}(\nu) \setminus V_{\text{far}}, \nu) \frac{\gamma - 1 - \hat{y}}{\gamma - 1}.$$

Then, using the case condition and simple algebra, we have

$$\begin{aligned} C_{\text{cls}}^{\text{max}}(\nu) &\leq D(\bar{N}_{\text{far}}(\nu) \setminus V_{\text{far}}, \nu) \\ &\leq C_{\text{far}}^{\text{avg}}(\nu) - \frac{\hat{y}\delta}{\gamma - 1 - \hat{y}} \leq C_{\text{far}}^{\text{avg}}(\nu) - \frac{\hat{y}\delta}{1 - \hat{y}}, \end{aligned} \quad (5.13)$$

where the last step follows from $1 < \gamma < 2$.

On the other hand, since K , V_{cls} , and V_{far} form a partition of $\bar{N}_{\text{cls}}(\kappa)$, we have

$C_{\text{cls}}^{\text{avg}}(\kappa) = (1 - \hat{y})D(K \cup V_{\text{cls}}, \kappa) + \hat{y}D(V_{\text{far}}, \kappa)$. Then using the definition of δ we obtain

$$D(K \cup V_{\text{cls}}, \kappa) = C_{\text{cls}}^{\text{avg}}(\kappa) + \frac{\hat{y}\delta}{1 - \hat{y}}. \quad (5.14)$$

Now we are essentially done. If there exists some $\mu \in V_{\text{cls}}$ such that $d_{\mu\kappa} \leq C_{\text{cls}}^{\text{avg}}(\kappa) +$

$\hat{y}\delta/(1 - \hat{y})$, then we have

$$\begin{aligned}
D(K, \nu) &\leq D(K, \kappa) + d_{\mu\kappa} + d_{\mu\nu} \\
&\leq C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{cls}}^{\text{avg}}(\kappa) + \frac{\hat{y}\delta}{1 - \hat{y}} + C_{\text{cls}}^{\text{max}}(\nu) \\
&\leq C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu),
\end{aligned}$$

where we used (5.13) in the last step. Otherwise, from (5.14), we must have $D(K, \kappa) \leq$

$C_{\text{cls}}^{\text{avg}}(\kappa) + \hat{y}\delta/(1 - \hat{y})$. Choosing any $\mu \in V_{\text{cls}}$, it follows that

$$\begin{aligned}
D(K, \nu) &\leq D(K, \kappa) + d_{\mu\kappa} + d_{\mu\nu} \\
&\leq C_{\text{cls}}^{\text{avg}}(\kappa) + \frac{\hat{y}\delta}{1 - \hat{y}} + C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{cls}}^{\text{max}}(\nu) \\
&\leq C_{\text{cls}}^{\text{avg}}(\kappa) + C_{\text{cls}}^{\text{max}}(\kappa) + C_{\text{far}}^{\text{avg}}(\nu),
\end{aligned}$$

again using (5.13) in the last step.

This concludes the proof of (5.10). As explained earlier, Lemma 23 follows. ■

Next, we derive some estimates for the expected cost of direct connections. The next technical lemma is a generalization of Lemma 17. In Lemma 17 we bound the expected distance to the closest open facility in $\overline{N}(\nu)$, conditioned on at least one facility in $\overline{N}(\nu)$ being open. The lemma below provides a similar estimate for an arbitrary set A of facilities in $\overline{N}(\nu)$, conditioned on that at least one facility in set A is open. Recall that $D(A, \nu) = \sum_{\mu \in A} d_{\mu\nu} \bar{y}_\mu / \sum_{\mu \in A} \bar{y}_\mu$ is the average distance from ν to a facility in A .

Lemma 24 *For any non-empty set $A \subseteq \overline{N}(\nu)$, let Λ_A^ν be the event that at least one facility in A is opened by Algorithm EBGs, and denote by $C_\nu(A)$ the random variable representing*

the distance from ν to the closest open facility in A . Then the expected distance from ν to the nearest open facility in A , conditioned on at least one facility in A being opened, is

$$\mathbb{E}[C_\nu(A) \mid \Lambda_A^\nu] \leq D(A, \nu).$$

Proof. The proof follows the same reasoning as the proof of Lemma 17, so we only sketch it here. We start with a similar grouping of facilities in A : for each primary demand κ , if $\bar{N}_{\text{cls}}(\kappa) \cap A \neq \emptyset$ then $\bar{N}_{\text{cls}}(\kappa) \cap A$ forms a group. Facilities in A that are not in a neighborhood of any primary demand form singleton groups. We denote these groups G_1, \dots, G_k . It is clear that the groups are disjoint because of (PD'.1). Denoting by $\bar{d}_s = D(G_s, \nu)$ the average distance from ν to a group G_s , we can assume that these groups are ordered so that $\bar{d}_1 \leq \dots \leq \bar{d}_k$.

Each group can have at most one facility open and the events representing opening of any two facilities that belong to different groups are independent. To estimate the distance from ν to the nearest open facility in A , we use an alternative random process to make connections, that is easier to analyze. Instead of connecting ν to the nearest open facility in A , we will choose the smallest s for which G_s has an open facility and connect ν to this facility. (Thus we selected an open facility with respect to the minimum \bar{d}_s , not the actual distance from ν to this facility.) This can only increase the expected connection cost, thus denoting $g_s = \sum_{\mu \in G_s} \gamma \bar{y}_\mu$ for all $s = 1, \dots, k$, and letting $\mathbb{P}[\Lambda_A^\nu]$ be the probability that

A has at least one facility open, we have

$$\mathbb{E}[C_\nu(A) \mid \Lambda_A^\nu] \leq \frac{1}{\mathbb{P}[\Lambda_A^\nu]} (\bar{d}_1 g_1 + \bar{d}_2 g_2 (1 - g_1) + \dots + \bar{d}_k g_k (1 - g_1) \dots (1 - g_{k-1})) \quad (5.15)$$

$$\begin{aligned} &\leq \frac{1}{\mathbb{P}[\Lambda_A^\nu]} \frac{\sum_{s=1}^k \bar{d}_s g_s}{\sum_{s=1}^k g_s} (1 - \prod_{s=1}^k (1 - g_s)) \\ &= \frac{\sum_{s=1}^k \bar{d}_s g_s}{\sum_{s=1}^k g_s} = \frac{\sum_{\mu \in A} d_{\mu\nu} \gamma \bar{y}_\mu}{\sum_{\mu \in A} \gamma \bar{y}_\mu} \\ &= \frac{\sum_{s=1}^k d_{\mu\nu} \bar{y}_\mu}{\sum_{\mu \in A} \bar{y}_\mu} = D(A, \nu). \end{aligned} \quad (5.16)$$

Inequality (5.16) follows from inequality (A.3) in A.2. The rest of the derivation follows from $\mathbb{P}[\Lambda_A^\nu] = 1 - \prod_{s=1}^k (1 - g_s)$, and the definition of \bar{d}_s , g_s and $D(A, \nu)$. ■

A consequence of Lemma 24 is the following corollary which bounds the other two expectations of C_ν , when at least one facility is opened in $\bar{N}_{\text{cls}}(\nu)$, and when no facility in $\bar{N}_{\text{cls}}(\nu)$ opens but a facility in $\bar{N}_{\text{far}}(\nu)$ is opened.

Corollary 25 (a) $\mathbb{E}[C_\nu \mid \Lambda_{\text{cls}}^\nu] \leq C_{\text{cls}}^{\text{avg}}(\nu)$, and (b) $\mathbb{E}[C_\nu \mid \Lambda^\nu \wedge \neg \Lambda_{\text{cls}}^\nu] \leq C_{\text{far}}^{\text{avg}}(\nu)$.

Proof. When there is an open facility in $\bar{N}_{\text{cls}}(\nu)$, the algorithm connect ν to the nearest open facility in $\bar{N}_{\text{cls}}(\nu)$. When no facility in $\bar{N}_{\text{cls}}(\nu)$ opens but some facility in $\bar{N}_{\text{far}}(\nu)$ opens, the algorithm connects ν to the nearest open facility in $\bar{N}_{\text{far}}(\nu)$. The rest of the proof follows from Lemma 24. By setting the set A in Lemma 24 to $\bar{N}_{\text{cls}}(\nu)$, we have

$$\mathbb{E}[C_\nu \mid \Lambda_{\text{cls}}^\nu] \leq D(\bar{N}_{\text{cls}}(\nu), \nu) = C_{\text{cls}}^{\text{avg}}(\nu),$$

proving part (a), and by setting the set A to $\bar{N}_{\text{far}}(\nu)$, we have

$$\mathbb{E}[C_\nu \mid \Lambda^\nu \wedge \neg \Lambda_{\text{cls}}^\nu] \leq D(\bar{N}_{\text{far}}(\nu), \nu) = C_{\text{far}}^{\text{avg}}(\nu),$$

which proves part (b). ■

Given the estimate on the three expected distances when ν connects to its close facility in $\overline{N}_{\text{cls}}(\nu)$ in (5.3), or its far facility in $\overline{N}_{\text{far}}(\nu)$ in (5.3), or its target facility $\phi(\kappa)$ in (5.10), the only missing pieces are estimates on the corresponding probabilities of each event, which we do in the next lemma. Once done, we shall put all pieces together and proving the desired inequality on $\mathbb{E}[C_\nu]$, that is (5.9).

The next Lemma bounds the probabilities for events that no facilities in $\overline{N}_{\text{cls}}(\nu)$ and $\overline{N}(\nu)$ are opened by the algorithm.

Lemma 26 (a) $\mathbb{P}[\neg\Lambda_{\text{cls}}^\nu] \leq 1/e$, and (b) $\mathbb{P}[\neg\Lambda^\nu] \leq 1/e^\gamma$.

Proof. (a) To estimate $\mathbb{P}[\neg\Lambda_{\text{cls}}^\nu]$, we again consider a grouping of facilities in $\overline{N}_{\text{cls}}(\nu)$, as in the proof of Lemma 24, according to the primary demand's close neighborhood that they fall in, with facilities not belonging to such neighborhoods forming their own singleton groups. As before, the groups are denoted G_1, \dots, G_k . It is easy to see that $\sum_{s=1}^k g_s = \sum_{\mu \in \overline{N}_{\text{cls}}(\nu)} \gamma \bar{y}_\mu = 1$. For any group G_s , the probability that a facility in this group opens is $\sum_{\mu \in G_s} \gamma \bar{y}_\mu = g_s$ because in the algorithm at most one facility in a group can be chosen and each is chosen with probability $\gamma \bar{y}_\mu$. Therefore the probability that no facility opens is $\prod_{s=1}^k (1 - g_s)$, which is at most $e^{-\sum_{s=1}^k g_s} = 1/e$. Therefore we have $\mathbb{P}[\neg\Lambda_A^\nu] \leq 1/e$.

(b) This proof is similar to the proof of (a). The probability $\mathbb{P}[\neg\Lambda^\nu]$ is at most $e^{-\sum_{s=1}^k g_s} = 1/e^\gamma$, because we now have $\sum_{s=1}^k g_s = \gamma \sum_{\mu \in \overline{N}(\nu)} \bar{y}_\mu = \gamma \cdot 1 = \gamma$. ■

We are now ready to bound the overall connection cost of Algorithm EBGs, namely inequality (5.9).

Lemma 27 *The expected connection of ν is*

$$\mathbb{E}[C_\nu] \leq C^{\text{avg}}(\nu) \cdot \max \left\{ \frac{1/e + 1/e^\gamma}{1 - 1/\gamma}, 1 + \frac{2}{e^\gamma} \right\}.$$

Proof. Recall that, to connect ν , the algorithm uses the closest facility in $\overline{N}_{\text{cls}}(\nu)$ if one is opened; otherwise it will try to connect ν to the closest facility in $\overline{N}_{\text{far}}(\nu)$. Failing that, it will connect ν to $\phi(\kappa)$, the sole facility open in the neighborhood of κ , the primary demand ν was assigned to. Given that, we estimate $\mathbb{E}[C_\nu]$ as follows:

$$\begin{aligned} \mathbb{E}[C_\nu] &= \mathbb{E}[C_\nu \mid \Lambda_{\text{cls}}^\nu] \cdot \mathbb{P}[\Lambda_{\text{cls}}^\nu] + \mathbb{E}[C_\nu \mid \Lambda^\nu \wedge \neg \Lambda_{\text{cls}}^\nu] \cdot \mathbb{P}[\Lambda^\nu \wedge \neg \Lambda_{\text{cls}}^\nu] \\ &\quad + \mathbb{E}[C_\nu \mid \neg \Lambda^\nu] \cdot \mathbb{P}[\neg \Lambda^\nu] \\ &\leq C_{\text{cls}}^{\text{avg}}(\nu) \cdot \mathbb{P}[\Lambda_{\text{cls}}^\nu] + C_{\text{far}}^{\text{avg}}(\nu) \cdot \mathbb{P}[\Lambda^\nu \wedge \neg \Lambda_{\text{cls}}^\nu] \\ &\quad + [C_{\text{cls}}^{\text{avg}}(\nu) + 2C_{\text{far}}^{\text{avg}}(\nu)] \cdot \mathbb{P}[\neg \Lambda^\nu] \\ &= [C_{\text{cls}}^{\text{avg}}(\nu) + C_{\text{far}}^{\text{avg}}(\nu)] \cdot \mathbb{P}[\neg \Lambda^\nu] + [C_{\text{far}}^{\text{avg}}(\nu) - C_{\text{cls}}^{\text{avg}}(\nu)] \cdot \mathbb{P}[\neg \Lambda_{\text{cls}}^\nu] + C_{\text{cls}}^{\text{avg}}(\nu) \\ &\leq [C_{\text{cls}}^{\text{avg}}(\nu) + C_{\text{far}}^{\text{avg}}(\nu)] \cdot \frac{1}{e^\gamma} + [C_{\text{far}}^{\text{avg}}(\nu) - C_{\text{cls}}^{\text{avg}}(\nu)] \cdot \frac{1}{e} + C_{\text{cls}}^{\text{avg}}(\nu) \\ &= \left(1 - \frac{1}{e} + \frac{1}{e^\gamma}\right) \cdot C_{\text{cls}}^{\text{avg}}(\nu) + \left(\frac{1}{e} + \frac{1}{e^\gamma}\right) \cdot C_{\text{far}}^{\text{avg}}(\nu). \end{aligned} \tag{5.18}$$

Inequality (5.17) follows from Corollary 25 and Lemma 23. Inequality (5.18) follows from Lemma 26 and $C_{\text{far}}^{\text{avg}}(\nu) - C_{\text{cls}}^{\text{avg}}(\nu) \geq 0$.

Now define $\rho = C_{\text{cls}}^{\text{avg}}(\nu)/C^{\text{avg}}(\nu)$. It is easy to see that ρ is between 0 and 1.

Continuing the above derivation, applying (5.8), we get

$$\begin{aligned} \mathbb{E}[C_\nu] &\leq C^{\text{avg}}(\nu) \cdot \left((1 - \rho) \frac{1/e + 1/e^\gamma}{1 - 1/\gamma} + \rho \left(1 + \frac{2}{e^\gamma}\right) \right) \\ &\leq C^{\text{avg}}(\nu) \cdot \max \left\{ \frac{1/e + 1/e^\gamma}{1 - 1/\gamma}, 1 + \frac{2}{e^\gamma} \right\}, \end{aligned}$$

and the proof is now complete. ■

With Lemma 27 proven, we are now ready to bound our total connection cost.

For any client j we have

$$\begin{aligned} \sum_{\nu \in j} C^{\text{avg}}(\nu) &= \sum_{\nu \in j} \sum_{\mu \in \mathbb{F}} d_{\mu\nu} \bar{x}_{\mu\nu} \\ &= \sum_{i \in \mathbb{F}} d_{ij} \sum_{\mu \in i} \sum_{\nu \in j} \bar{x}_{\mu\nu} = \sum_{i \in \mathbb{F}} d_{ij} x_{ij}^* = C_j^*. \end{aligned}$$

Summing over all clients j we obtain that the total expected connection cost is

$$\mathbb{E}[C_{\text{EBGS}}] \leq C^* \max \left\{ \frac{1/e + 1/e^\gamma}{1 - 1/\gamma}, 1 + \frac{2}{e^\gamma} \right\}.$$

Recall that the expected facility cost is bounded by γF^* , as argued earlier. Hence the total expected cost is bounded by $\max\{\gamma, \frac{1/e + 1/e^\gamma}{1 - 1/\gamma}, 1 + \frac{2}{e^\gamma}\} \cdot \text{LP}^*$. Picking $\gamma = 1.575$ we obtain the desired ratio.

Theorem 28 *Algorithm EBGs is a 1.575-approximation algorithm for FTFP.*

Chapter 6

Primal-dual Algorithms

In this chapter we present results and discuss combinatorial algorithms to the FTFP problem. These approaches, although employ the Linear Program in guiding the algorithm and deriving approximation ratio, the use of LP is implicit. Moreover, the algorithms do not require solving the LP and having access to the fractional optimal solutions. Two notable such approaches are primal-dual and dual-fitting. In primal-dual algorithms, we start with a feasible dual solution, usually with all dual variables set to zero, then we raise a subset of dual variables and update the corresponding primal variables accordingly. At any time, we keep the dual solution feasible and we stop when the primal solution becomes feasible. The approximation ratio is derived by a relaxed version of the complementary slackness conditions.

Another approach, dual-fitting, starts with an empty dual solution as well, and raise a subset of dual variables in each iteration, updating corresponding primal variables and stop when the primal solution is feasible. The difference is that in dual-fitting, the

dual solution may not be feasible, and we require the cost of the primal solution bounded by the value of the possibly infeasible dual solution. The next step is to find the smallest possible number γ , which may depend on the input size, such that the dual solution, when divided by γ , becomes feasible. It is easy to see that γ provides an upper bound on the approximation ratio, because the value of a feasible dual solution is a lower bound on the value of an optimal primal solution.

Jain and Vazirani [18] designed a primal-dual algorithm, which we call the JV algorithm, for the UFL problem. Recall that in the UFL, all demands $r_j = 1$. In the JV algorithm, every client j has a number α_j associated. All α_j start with zero. The algorithm has two phases. The first phase runs in iterations. In each iteration, all α_j that were not temporarily connected are raised uniformly. The contribution from a client j to a facility i is $\max\{0, \alpha_j - d_{ij}\}$. Whenever a facility received enough total contribution, that is $\sum_{j \in C} (\alpha_j - d_{ij})_+ = f_i$, then i is temporarily open and all clients with $\alpha_j \geq d_{ij}$ temporarily connect to i . The facility i is called the witness of the client j . The first phase concludes when all clients are temporarily connected. In the second phase, we construct an auxiliary graph with nodes being temporarily open facilities and two nodes are connected by an edge if there exists some client j that contributes to both of them, that is, $\alpha_j > d_{i_1j}$ and $\alpha_j > d_{i_2j}$. We then pick a maximal independent set in the auxiliary graph as the set of facilities to open. For connections, if a client j has an open facility i with $d_{ij} \leq \alpha_j$, then it connects to that facility. Otherwise, there exists some temporarily open facility i such that $\alpha_j \geq d_{ij}$. Since i is not open, there must exist some facility i' that is open and some client j' such that $\alpha_{j'} > d_{i'j}$ and $\alpha_{j'} > d_{ij'}$. It follows that $\alpha_{j'} \leq \min\{t(i), t(i')\}$

where $t(i)$ is the time that facility i is temporarily open. The reason is, if $\alpha_{j'} > t(i)$, then it would have temporarily connected to facility i earlier so its $\alpha_{j'}$ value would have been smaller. On the other hand, since facility i is the witness of client j , we have $t(i) \leq \alpha_j$ ¹. Therefore we have $\alpha_j \geq \alpha_{j'} \geq d_{i'j'}, d_{ij}$. In addition, we also have $\alpha_j \geq d_{ij}$. Hence $d_{i'j} \leq d_{i'j'} + d_{ij'} + d_{ij} \leq \alpha_{j'} + \alpha_{j'} + \alpha_j \leq 3\alpha_j$. If we define $\beta_{ij} = 0$ if client j does not contribute to any facility i and $\beta_{ij} = \alpha_j - d_{ij}$ if client j connects directly to facility i . Now we estimate the total cost of this solution. For facility cost we have $\sum_{i,j} \beta_{ij}$, and for connection cost, if a client j is directly connected, then its $d_{ij} \leq \alpha_j - \beta_{ij}$, otherwise it is $d_{ij} \leq 3\alpha_j$. The total cost is hence no more than $3 \sum_{j \in \mathbb{C}} \alpha_j$. Since $\{\alpha_j\}$ form a feasible dual solution, we have the optimal solution value is no less than $\sum_{j \in \mathbb{C}} \alpha_j$. Therefore, we have our solution costs no more than 3 times the cost of an optimal solution.

The fault-tolerant facility location problem (FTFL) was introduced by Jain and Vazirani [19] primarily to demonstrate that their primal-dual algorithm on UFL can be applied to a more general problem, where clients could have demand more than 1, and each facility could be open or close. A client j with demand r_j needs to be connected to r_j different facilities. The primal-dual algorithm they provide gives a ratio of $3 \ln \max_j r_j$. Subsequent attempts on adapting either the primal-dual approach or the dual-fitting approach to FTFL with a sub-logarithmic approximation ratio were not successful, although for the uniform demands case, that is, when all r_j are equal, Adrian Bumb [2] demonstrated that the JV algorithm [18] for UFL can be adapted to obtain the same ratio as for UFL. On a separate paper, Swamy and Shmoys [25] showed that a greedy algorithm analyzed

¹ $t(i) < \alpha_j$ is possible if facility i is temporarily open and later j has $\alpha_j = d_{ij}$ to temporarily connect to facility i .

using dual-fitting can be shown to have a ratio of 1.52. For the non-uniform demand case, the best known result is an $O(\log n)$ -approximation.

For our problem, FTFP, we have seen in earlier chapters that it can be approximated with the same ratio as UFL when LP-rounding is used. However, the attempt to obtain a sub-logarithmic approximation ratio on FTFP using the primal-dual algorithm or the dual-fitting algorithm were not successful. On the positive side, we derive a weak result that the greedy algorithm does give a $O(\log n)$ ratio for FTFP. We remark here that the $O(\log n)$ ratio does not even use the triangle inequality. On the negative side, we provide an example showing that the greedy algorithm with dual-fitting analysis can at best give a ratio of $O(\log n / \log \log n)$ under a very reasonable assumption, which we call *local-charging* assumption. Here we have $n = |\mathbb{C}|$.

In the following we first describe the greedy algorithm and its analysis. We show that the greedy algorithm is $O(\log n)$ approximation using dual-fitting analysis. Then we present our example showing a lower bound on the dual-fitting analysis on the greedy algorithm. We conclude this chapter with some possible approach to obtain sub-logarithmic approximation results.

6.1 The Greedy algorithm with $O(\log n)$ Ratio

In this section we show that the greedy algorithm which repeatedly picking the best star (the one with minimum average cost) gives an approximation ratio of $H_n = \ln(n)$ where $n = |\mathbb{C}|$ is the number of clients. A star is a site i and a subset of clients C' . The cost of such a star S is $c(S) = f_i + \sum_{j \in C'} d_{ij}$, and the average cost of S is $c(S)/|C'|$. Call a client

j fully-satisfied if j has made r_j connections. Let U be the set of not fully-satisfied clients. While not all clients fully-satisfied, the algorithm picks a star $S = (i, C')$ with $C' \subseteq U$, and open one facility at site i . Each client in C' then makes one more connection with site i . The algorithm terminates when all clients are fully-satisfied.

When we run the greedy algorithm, for every client j , we associate each demand of j with a number α_j^l , which is the average cost of the star when l^{th} demand of j is connected. Now we let $\alpha_j = \alpha_j^{r_j}$, that is, take α_j to be the finishing α_j^l , and order clients by increasing α_j . That is,

$$\alpha_1 \leq \alpha_2 \leq \dots \leq \alpha_n$$

Due to the algorithm, for every $j = 1, \dots, n$, we have

$$\sum_{l=j}^n (\alpha_j - d_{il})_+ \leq f_i$$

for every site i . The reason is that, when the last demand of j is connected, all clients $j + 1, \dots, n$ are still active so their total contribution cannot exceed f_i .

Now we take a closer look at the numbers $\{\alpha_j\}$. We know that the algorithm's total cost is exactly $\sum_{j=1}^n \sum_{l=1}^{r_j} \alpha_j^l$, which is no more than $\sum_{j=1}^n r_j \alpha_j$ since we take α_j to be $\alpha_j^{r_j}$. Now if we can show that $\sum_{j=1}^n r_j \alpha_j$ is no more than $\gamma \cdot \text{OPT}$, where OPT is the cost of an integral optimal solution to the given FTFP instance, then we claim our algorithm returns an integral solution within a factor of γ .

We show that $\sum_{j=1}^n r_j \alpha_j$ is within a factor of γ from OPT by showing that $\{\alpha_j/\gamma\}$ is a feasible dual solution to the following program, which is the dual program of the primal

LP for FTFP.

$$\begin{aligned} & \max \sum_j r_j \alpha_j \\ \text{subject to: } & \sum_{j=1}^n (\alpha_j - d_{ij})_+ \leq f_i \text{ for every facility } i \end{aligned}$$

To find the minimum γ that would shrink $\{\alpha_j\}$ to a feasible dual solution, we need to find a worst case instance to maximize γ , also it is clear that the worst case instance must contain a star whose feasibility requirement would achieves the value of γ , and this star would be the worst star in that instance.

As a first step we can drop the $\max\{0, \cdot\}$, because we can always find a new star by dropping those j with $\alpha_j - d_{ij}$ term negative, and that new star would still be a worst case star. Suppose a worst case star has k clients, and is with facility i , then we have

$$\sum_{j=1}^k \alpha_j - d_{ij} \leq f_i$$

Here we rename clients in the new star to be $1, \dots, k$, although among them, they are still ordered by their α_j .

Now our goal is to find a supremum of the following program:

$$\begin{aligned} & \max \frac{\sum_{j=1}^k \alpha_j}{f_i + \sum_{j=1}^k d_{ij}} \\ \text{subject to: } & \sum_{l=j}^k (\alpha_l - d_{il})_+ \leq f_i \text{ for } j = 1, \dots, k \end{aligned}$$

Since we are dealing with a particular star, we can abstract away i , to obtain the following program:

$$\begin{aligned} & \max \frac{\sum_{j=1}^k \alpha_j}{f + \sum_{j=1}^k d_j} \\ & \text{subject to: } \sum_{l=j}^k (\alpha_j - d_l)_+ \leq f \text{ for } j = 1, \dots, n \end{aligned} \tag{6.1}$$

Now we claim we can drop the $\max\{0, \cdot\}$ operator because this would relax the constraint in (6.1) and can only make objective value larger (since we are maximizing), so the real optimal is upper bounded by the relaxed optimal. This allows us to work on

$$\begin{aligned} & \max \frac{\sum_{j=1}^k \alpha_j}{f + \sum_{j=1}^k d_j} \\ & \text{subject to: } \sum_{l=j}^k (\alpha_j - d_l) \leq f \text{ for } j = 1, \dots, n \end{aligned} \tag{6.2}$$

For each $j = 1, \dots, n$, the constraint above simply can be rewritten as

$$(k - j + 1)\alpha_j \leq f + \sum_{l=j}^k d_l \leq f + \sum_{l=1}^k d_l. \tag{6.3}$$

The first inequality is a rewrite of the constraint in (6.2) and the second is straightforward.

Therefore we have $\alpha_j \leq (1/(k - j + 1))(f + \sum_{l=j}^k d_l)$, and it easily follows that

$$\sum_{j=1}^n \alpha_j \leq (1/k + 1/(k-1) + \dots + 1) = H_k \leq H_n = \ln(n) \tag{6.4}$$

6.2 An Example Showing the Difficulty in Obtaining $O(1)$ Ratio

For FTFP, the greedy algorithm that repeatedly picks the best star until all clients have all demands satisfied can be implemented in polynomial time. In Section 6.1 we show that this algorithm is H_n -approximation where $n = |\mathcal{C}|$ is the number of clients. Since the same greedy algorithm is shown to have constant approximation ratio for UFL [21], a natural question to ask is whether greedy can be shown to have $O(1)$ approximation ratio. Here we give an argument that hints a negative answer.

We assume the greedy algorithm is analyzed using the dual-fitting technique, which associates with every client j with a number α_j , interpreted as a dual solution to the LP (3.2). However, the dual solution in general may not be feasible. The dual-fitting technique aims at finding a smallest possible number γ such that after the dual solution $\{\alpha_j\}$ is shrunk (divided) by γ , all dual constraints are satisfied. That is

$$\sum_{j \in \mathcal{C}} (\alpha_j / \gamma - d_{ij})_+ \leq f_i \quad \text{for all } i \in \mathcal{F}.$$

That γ is taken as the approximation ratio.

In the greedy algorithm, a star with minimum average cost is picked at each iteration and each member client of that star then gets one more connection. It is not specified by the algorithm how we distribute the cost of f_i into member clients, which is part of the analysis. Nonetheless we assume that the cost of f_i is distributed among members only, and not to clients outside this star. We call this *local charging* assumption. Our second assumption is that the proposed dual solution α_j , is taken as the average of

individual α_j^l for each of the l^{th} demand of client j , with $l = 1, \dots, r_j$. Suppose the l^{th} demand of j is satisfied while j is in a star with facility i , then $\alpha_j^l = d_{ij} + f_i^{j,l}$, where $f_i^{j,l}$ is the portion of f_i attributed to j in the analysis. Notice that taking the average implies the α_j values thus computed results in $\sum_{j \in C} r_j \alpha_j$ equal to the cost of the integral solution by the greedy algorithm.

Our example has one site and k groups of clients, a figure is given in Figure 6.1. Opening one facility at that site costs f_1 . The first group has n_1 clients each with demand r_1 , all at distance $d_1 = 0$ from f_1 . The other groups are listed below:

$$d_1 = 0$$

$$d_2 = \frac{f_1}{n_1}$$

$$d_3 = f_1/n_2 + d_2 = f_1/n_2 + f_1/n_1 = f_1\left(\frac{1}{n_2} + \frac{1}{n_1}\right)$$

$$\dots$$

$$d_k = f_1/n_{k-1} + d_{k-1} = f_1\left(\frac{1}{n_{k-1}} + \dots + \frac{1}{n_1}\right)$$

For the numbers, we need $r_1 \ll r_2 \ll \dots \ll r_k$, and $n_1 = u^{k-1}, n_2 = u^{k-2}, \dots, n_k = u^0 = 1$ for some number u (Actually we take $u = k$, this choice may not be the best possible).

Call a star with facility cost zero *trivial*. It is *non-trivial* if the facility has non-zero cost. Now the greedy execution goes like this: The first non-trivial star (with r_1 replica) is (f_1, n_1) . Then we have a trivial star of zero cost facility and all n_2 clients in group 2 for r_1 replica. The second non-trivial star (with r_2 replica) is (f_1, n_2) . Notice that $r_2 \gg r_1$. The r_1 replica of trivial star with group 2 satisfy r_1 demand of the n_2 group. After that the

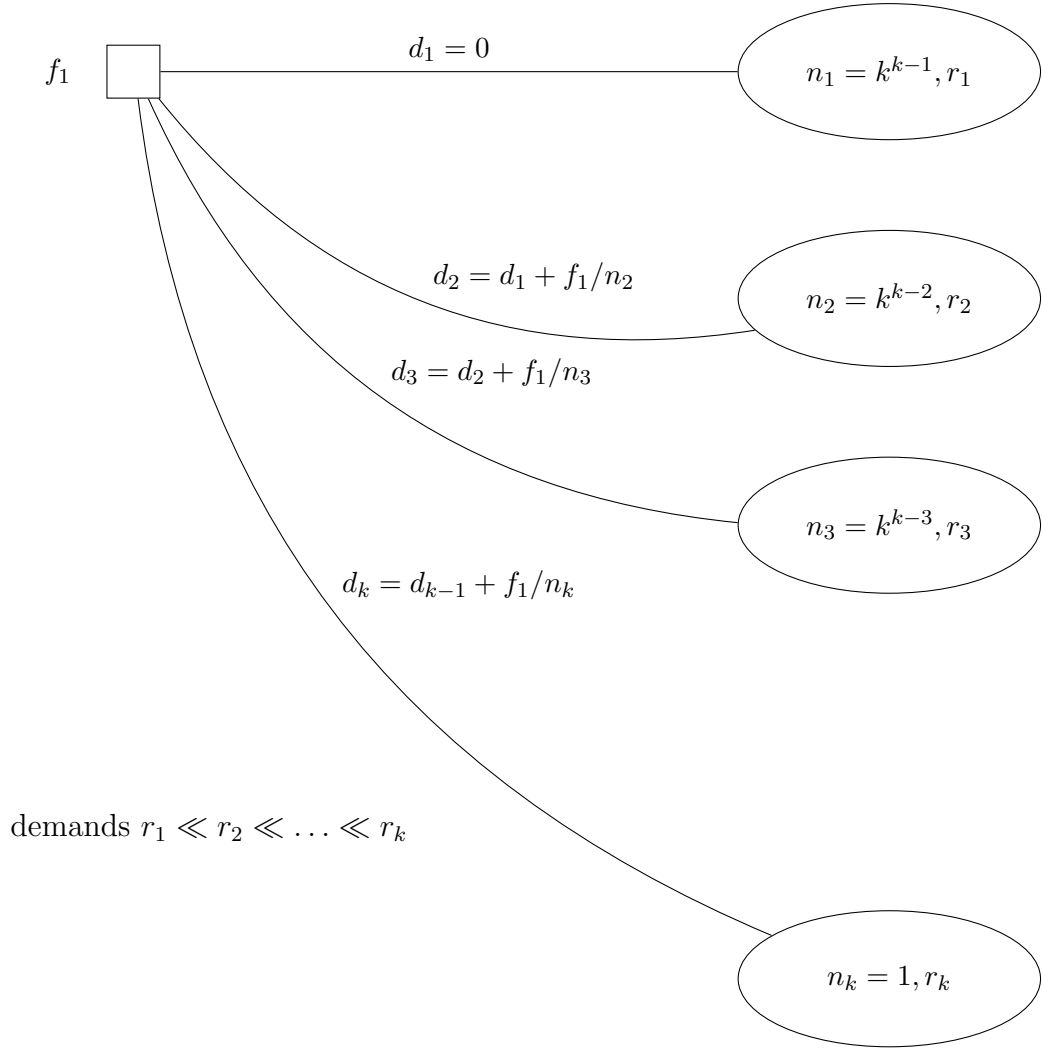


Figure 6.1: An example showing the greedy algorithm for FTFP, analyzed using dual-fitting, could give a solution with cost $\Omega(\log n / \log \log n)$ from the optimal value, assuming facility cost can only be charged to clients within the star.

n_2 group clients each has residual demand $r_2 - r_1 = r_2$. The process repeats until the k^{th} group finishes with r_k new facilities.

According to our local charging assumption, we have $\alpha_1 = f_1$, now defined as the total dual value of clients in group n_1 , regardless how the analysis would distribute within that group. Similarly $\alpha_2 = f_1 + n_2 d_2$, and so on. Substituting in the numbers, we have

$$\alpha_1 = f_1$$

$$\alpha_2 = f_1 + n_2 d_2 = f_1 + f_1/n_1 \cdot n_2 = f_1(1 + n_2/n_1)$$

$$\alpha_3 = f_1 + n_3 d_3 = f_1 + f_1(\frac{1}{n_2} + \frac{1}{n_1})n_3 = f_1(1 + \frac{n_3}{n_2} + \frac{n_3}{n_1})$$

\dots

$$\alpha_k = f_1 + n_k d_k = f_1 + f_1 n_k (\frac{1}{n_{k-1}} + \dots \frac{1}{n_1})$$

Notice that $r_1 \ll r_2 \ll \dots \ll r_k$ implies α_j is decided by the max among α_j^l .

Now back to the dual constraint, it requires that the shrinking factor γ needs to satisfy the following inequality:

$$\frac{\alpha_1}{\gamma} - d_1 + \frac{\alpha_2}{\gamma} - d_2 + \dots + \frac{\alpha_k}{\gamma} - d_k \leq f_1. \quad (6.5)$$

Substitute in the α_j values derived above, we have

$$\begin{aligned}
\gamma &\geq (\sum_{j=1}^k \alpha_j) / (f_1 + \sum_{j=1}^k d_j) \\
&\geq \frac{f_1 + n_1 d_1 + f_1 + n_2 d_2 + f_1 + n_3 d_3 + \dots + f_1 + n_k d_k}{f_1 + n_1 d_1 + n_2 d_2 + \dots + n_k d_k} \\
&= 1 + (k-1)f_1 / (f_1 + n_1 d_1 + n_2 d_2 + \dots + n_k d_k) \\
&= 1 + (k-1)f_1 / \left(f_1 + n_2 f_1 / n_1 + \dots + n_k f_1 \left(\frac{1}{n_{k-1}} + \frac{1}{n_{k-2}} + \dots + \frac{1}{n_1} \right) \right) \\
&= 1 + (k-1) / \left(1 + n_2 / n_1 + \dots + n_k \left(\frac{1}{n_{k-1}} + \frac{1}{n_{k-2}} + \dots + \frac{1}{n_1} \right) \right) \\
&= 1 + (k-1) / \left(1 + 1/u + \dots + \left(\frac{1}{u} + \dots + \frac{1}{u^{k-1}} \right) \right) \\
&= 1 + (k-1) / \left(1 + k/u + (k-1)/u^2 + \dots + 1/u^{k-1} \right) \\
&\geq 1 + (k-1) / \left(1 + k/u + k/u^2 + \dots + k/u^{k-1} \right) \\
&= 1 + (k-1) / \left(1 + 1 + 1/k + \dots + 1/k^{k-2} \right) \\
&\approx k/2
\end{aligned}$$

So for k groups we can force a shrinking factor γ as big as $k/2$. Recall that we have greedy being no more than H_n -approximation. Is that a contradiction? No, because we have the number of clients $n = k^{k-1} + k^{k-2} + \dots + 1 = k^k$, so $k = O(\log n / \log \log n)$. Therefore, the example shows that dual fitting with local charging cannot hope to get $O(\log n / \log \log n)$ ratio or better.

Remark Notice this example is similar in spirit to the $\Omega(\log n / \log \log n)$ example for Hochbaum's algorithm for UFL, constructed by Mahdian *et al.* [17].

Chapter 7

Conclusion

In this dissertation we studied the fault-tolerant facility placement problem (FTFP), a generalization of the well-known uncapacitated facility location problem (UFL). We demonstrated that the known LP-rounding algorithms for UFL can be adapted to FTFP while preserving the approximation ratio. To accomplish this reduction, we developed two techniques, namely demand reduction and adaptive partition, which could be of more general interest. Our results demonstrated that FTFP seems easier to approximate, compared to FTFL.

We also studied the primal-dual and dual-fitting approach, and provided a possible explanation on the difficulty to obtain a constant approximation ratio using those techniques.

We hope our work in this dissertation will help other researchers interested in the fault-tolerant variant of the facility location problems to develop more insight into the difficulty and possible solutions when clients demand more than one facility and we still need to keep total cost under control.

In anticipating future research, we tend to agree with the authors, Byrka *et al.*, with their remark on UFL and FTFL, that both problems are likely to have approximation algorithms with ratio matching the 1.463 lower bound. And from our demand reduction technique, it is almost surely that FTFP shall have a 1.463-approximation algorithm, provided that FTFL can be approximated to meet the lower bound.

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Appendix A

Technical Background

A.1 Linear Programming and Integer Programming

In this section we give a short introduction on Linear Programming and Integer Programming with an emphasis on their applicability in design and analysis of approximation algorithms for NP-optimization problems.

Most NP-optimization problems have a natural integer program in which we use variables to represent parameters in the solution we seek, and write the constraints imposed by feasibility of the solution, the objective function is obtained by the cost function of the solution, specified by the problem. For example, in the Vertex Cover problem, we are given a graph $G = (V, E)$ and we are to find a subset W of V , such that every edge $e \in E$ has at least one endpoint in W , and we want the set W to have minimum size. To formulate this problem as an integer program, we use $x_v \in \{0, 1\}$ to denote whether a node $v \in V$ is in W or not. The constraint is that for every edge $e = (u, v)$, we have $x_u + x_v \geq 1$. The objective

is to minimize $\sum_{v \in V} x_v$. The integer program for Vertex Cover is written as

$$\begin{aligned} & \text{minimize } \sum_{v \in V} x_v \\ & \text{subject to } x_u + x_v \geq 1 \quad \forall (u, v) \in E \\ & x_v \in \{0, 1\} \quad \forall v \in V \end{aligned}$$

In general an integer program cannot be solved exactly in polynomial time, as integer programming is **NP**-hard. However, if we relax the integral constraint and allow the variables to take fractional value, we then obtain a Linear Program (LP) and LP is polynomially solvable, for example, using the ellipsoid method or the interior point method. Thus we can first solve the LP optimally, obtaining a fractional optimal solution to the LP. The value of the fractional optimal solution is then a lower bound on the value of the optimal integral solution, assuming a minimization problem. Our next step is then to round the fractional solution appropriately, so that we maintain the feasibility while keeping the cost from increasing too much. The exact rounding procedure is problem specific and we shall not delve into the details here. The rounding relevant to the FTFP problem in this thesis is presented in detail in Chapter 5.

We now give a brief introduction of linear programming, see [10] for an introductory

book on this topic. A general Linear Program can be written as

$$\begin{aligned}
& \text{minimize} && \sum_{j=1}^n c_j x_j && (A.1) \\
& \text{subject to} && \sum_{j=1}^n a_{ij} x_j \geq b_i, && \text{for } i = 1, \dots, m \\
& && x_j \geq 0 && \text{for } j = 1, \dots, n
\end{aligned}$$

For the LP above, we can take its dual as

$$\begin{aligned}
& \text{maximize} && \sum_{i=1}^m b_i y_i && (A.2) \\
& \text{subject to} && \sum_{i=1}^m a_{ij} y_i \leq c_j && \text{for } j = 1, \dots, n \\
& && y_i \geq 0 && \text{for } i = 1, \dots, m
\end{aligned}$$

The LP (A.1) is called the primal program and the LP (A.2) is called the dual program.

The weak duality theorem tells us that, for every feasible solution \mathbf{x} for the primal (A.1)

and \mathbf{y} for the dual (A.2), we have that $\mathbf{c}^T \mathbf{x} \geq \mathbf{b}^T \mathbf{y}$. The strong duality theorem tells

us that if both the primal (A.1) and the dual (A.2) are feasible, then both of them have

optimal solution \mathbf{x}^* and \mathbf{y}^* and their objective function values equal, that is $\mathbf{c}^T \mathbf{x}^* = \mathbf{b}^T \mathbf{y}^*$.

Moreover, the complementary slackness conditions assert that two feasible solutions \mathbf{x} and \mathbf{y}

are both optimal to LP (A.1) and (A.2) respectively, if and only if, for every primal variable

x_j , either $x_j = 0$ or the corresponding constraint in the dual is tight, that is $\sum_{i=1}^m a_{ij} y_i = c_j$.

And for every dual variable y_i , either $y_i = 0$ or the corresponding constraint in the primal

is tight, that is $\sum_{j=1}^n a_{ij}x_j = b_i$. The complementary slackness conditions provide a simple way to validate the optimality when one is presented with two solutions, proposed to be optimal for the primal and dual program respectively.

The complementary slackness conditions play a crucial role in the design and analysis of approximation algorithms. For example, suppose we have an algorithm that computes a feasible integral solution \mathbf{x} to the primal program (A.1) and a feasible integral solution to the dual program (A.2). Moreover, we know that the two solutions satisfy a relaxed version of the complementary slackness conditions: for some numbers α and β , we have

$$\begin{aligned} \text{either } y_i = 0 \quad \text{or} \quad b_i \leq \sum_j a_{ij}x_j \leq \alpha b_i \quad & \text{for } i = 1, \dots, m. \\ \text{either } x_j = 0 \quad \text{or} \quad \beta c_j \leq \sum_i a_{ij}y_i \leq c_j \quad & \text{for } j = 1, \dots, n. \end{aligned}$$

Then the integral solution \mathbf{x} has cost no more than α/β times the optimal value. In particular, we have $\sum_j c_j x_j \leq \alpha/\beta \sum_i b_i y_i$ and the value for a feasible dual solution, namely $\sum_i b_i y_i$, is a lower bound on the optimal value of the primal program.

As an application of the complementary slackness conditions, we look at their use in the design and analysis of algorithms for the facility location problems. Recall that we define the neighborhood $N(j)$ of a client j as the set of facilities with $x_{ij}^* > 0$, where $\mathbf{x}^*, \mathbf{y}^*$ is some fractional optimal fractional solution and $\boldsymbol{\alpha}^*, \boldsymbol{\beta}^*$ is some optimal fractional dual solution. The complementary slackness conditions provide an upper bound on the maximum distance from a facility $i \in N(j)$ to a client j , since one dual constraint says

$\alpha_j - \beta_{ij} \leq d_{ij}$ and if the primal solution has $x_{ij}^* > 0$, then the inequality is actually an equality and we have $\alpha_j^* - \beta_{ij}^* = d_{ij}$. Together with $\beta_{ij}^* \geq 0$, we have $\alpha_j^* \geq d_{ij}$ for every i such that $x_{ij}^* > 0$.

The idea of using relaxed complementary slackness conditions in designing approximation algorithms for the uncapacitated facility location problem is demonstrated by Jain and Vazirani [18]. They proposed an algorithm that outputs an integral solution (\mathbf{x}, \mathbf{y}) to the primal program (3.1) and a feasible (possibly fractional) solution $(\boldsymbol{\alpha}, \boldsymbol{\beta})$ to the dual program (3.2)¹. Moreover, the two solutions satisfy the conditions that

$$\text{either } \sum_j \beta_{ij} = f_i \quad \text{or} \quad y_i = 0.$$

$$\text{either } 1/3 d_{ij} \leq \alpha_j - \beta_{ij} \leq d_{ij} \quad \text{or} \quad x_{ij} = 0.$$

The solution (\mathbf{x}, \mathbf{y}) then is an 3-approximation to the optimal solution.

A.2 Proof of Inequality (5.3)

In Sections 5.2 and 5.3 we use the following inequality

$$\begin{aligned} & \bar{d}_1 g_1 + \bar{d}_2 g_2 (1 - g_1) + \dots + \bar{d}_k g_k (1 - g_1)(1 - g_2) \dots (1 - g_{k-1}) \\ & \leq \frac{1}{\sum_{s=1}^k g_s} \left(\sum_{s=1}^k \bar{d}_s g_s \right) \left(\sum_{t=1}^k g_t \prod_{z=1}^{t-1} (1 - g_z) \right). \end{aligned} \tag{A.3}$$

for $0 < \bar{d}_1 \leq \bar{d}_2 \leq \dots \leq \bar{d}_k$, and $0 < g_1, \dots, g_s \leq 1$.

We give here a new proof of this inequality, much simpler than the existing proof

¹For the uncapacitated facility location problem we have all $r_j = 1$ for $j \in \mathbb{C}$.

in [8], and also simpler than the argument by Sviridenko [24]. We derive this inequality from the following generalized version of the Chebyshev Sum Inequality:

$$\sum_i p_i \sum_j p_j a_j b_j \leq \sum_i p_i a_i \sum_j p_j b_j, \quad (\text{A.4})$$

where each summation runs from 1 to l and the sequences (a_i) , (b_i) and (p_i) satisfy the following conditions: $p_i \geq 0, a_i \geq 0, b_i \geq 0$ for all i , $a_1 \leq a_2 \leq \dots \leq a_l$, and $b_1 \geq b_2 \geq \dots \geq b_l$.

Given inequality (A.4), we can obtain our inequality (A.3) by simple substitution

$$p_i \leftarrow g_i, a_i \leftarrow \bar{d}_i, b_i \leftarrow \prod_{s=1}^{i-1} (1 - g_s),$$

for $i = 1, \dots, k$.

For the sake of completeness, we include the proof of inequality (A.4), due to Hardy, Littlewood and Polya [15]. The idea is to evaluate the following sum:

$$\begin{aligned} S &= \sum_i p_i \sum_j p_j a_j b_j - \sum_i p_i a_i \sum_j p_j b_j \\ &= \sum_i \sum_j p_i p_j a_j b_j - \sum_i \sum_j p_i a_i p_j b_j \\ &= \sum_j \sum_i p_j p_i a_i b_i - \sum_j \sum_i p_j a_j p_i b_i \\ &= \frac{1}{2} \cdot \sum_i \sum_j (p_i p_j a_j b_j - p_i a_i p_j b_j + p_j p_i a_i b_i - p_j a_j p_i b_i) \\ &= \frac{1}{2} \cdot \sum_i \sum_j p_i p_j (a_i - a_j)(b_i - b_j) \leq 0. \end{aligned}$$

The last inequality holds because $(a_i - a_j)(b_i - b_j) \leq 0$, since the sequences (a_i) and (b_i) are ordered oppositely.