

EECS598-001

# Approximation Algorithms & Hardness of Approximation

Pingbang Hu

September 22, 2022

## Abstract

This is an advanced graduate-level algorithm course taught in University of Michigan by [Euiwoong Lee](#). Topics include both approximation algorithms like covering, clustering, network design, and constraint satisfaction problems (the first half), and also the hardness of approximation algorithms (the second half).

The first half of the course is classical and well-studied, and we'll use Williamson and Shmoys [[WS11](#)], Vazirani [[Vaz02](#)] as our reference. The second half of the course is still developing, and we'll look into papers by Barak and Steurer [[BS14](#)], O'donnell [[ODo21](#)], etc.

This course is taken in Fall 2022, and the date on the covering page is the last updated time.

# Contents

<b>1</b>	<b>Introduction</b>	<b>2</b>
1.1	Computational Problem . . . . .	2
1.2	Efficient Algorithms . . . . .	2
1.3	Approximation Algorithms . . . . .	3
1.4	Hardness . . . . .	4
<b>2</b>	<b>Covering</b>	<b>5</b>
2.1	Set Cover . . . . .	5
2.2	Greedy Method . . . . .	6
2.3	Linear Programming Rounding . . . . .	7
2.4	Covering-Packing Duality . . . . .	10
2.5	Primal-Dual Method . . . . .	11
2.6	Feedback Vertex Set . . . . .	12
<b>3</b>	<b>Clustering</b>	<b>18</b>
3.1	Clustering . . . . .	18
3.2	Facility Location . . . . .	18
3.3	$k$ -Median . . . . .	25

# Chapter 1

## Introduction

### Lecture 1: Overview, Set Cover

#### 1.1 Computational Problem

29 Aug. 10:30

We're interested in the following optimization problem: Given a problem with an input, we want to either maximize or minimize some objectives. This suggests the following definition.

**Definition 1.1.1 (Computational problem).** A *computational problem*  $P$  is a function from input  $I$  to  $(X, f)$ , where  $X$  is the feasible set of  $I$  and  $f$  is the objective function.

We see that by replacing  $f$  with  $-f$ , we can unify the notion and only consider either minimization or maximization, but we will not bother to do this.

**Example ( $s$ - $t$  shortest path).** The  $s$ - $t$  shortest path problem  $P$  can be formalized as follows. Given input  $I$ , it defines

- Input: Graph  $\mathcal{G} = (\mathcal{V}, \mathcal{E})$  and two vertices  $s, t \in \mathcal{V}$ .
- Feasible set:  $X = \{\text{set of all (simple) paths } s \text{ to } t\}$ .
- Objective function:  $f: X \rightarrow \mathbb{R}$  where  $f(x) = \text{length}(x)$  ( $\#$  of edges of  $x$ ).

The output of  $P$  should be some  $x \in X$  (i.e., some valid  $s$ - $t$  paths) such that it minimizes  $f(x)$ .

We see that the [computational problem](#) we focus on is an optimization problem, and more specifically, we're interested in [combinatorial optimization](#).

**Definition 1.1.2 (Combinatorial optimization).** A *combinatorial optimization* problem is a [problem](#) where the feasible set  $X$  is a finite set.

**Example ( $s$ - $t$  shortest path).** The  $s$ - $t$  shortest path problem is a [combinatorial optimization](#) problem since given a graph  $\mathcal{G}$  with  $n = |\mathcal{V}|$ ,  $m = |\mathcal{E}|$ , there are at most  $n!$  different paths, i.e.,  $|X| \leq n! < \infty$ .

**Note.** We'll also look into some continuous optimization problem, where  $X$  is now infinite (or even uncountable). For example, find  $x \in \mathbb{R}$  that minimizes  $f(x) = x^2 + 2x + 1$ . In this case,  $X = \mathbb{R}$  which is uncountable (hence infinite).

#### 1.2 Efficient Algorithms

Given a [problem](#)  $P$ , we want to solve it fast with [algorithms](#). Before we characterize the speed of an [algorithm](#), we should first define what exactly an algorithm is.

**Definition 1.2.1 (Algorithm).** Given a **problem**  $P$  and input  $I$  (which defines  $X$  and  $f$ ), an *algorithm*  $A$  outputs solution  $y = A(I)$  such that  $y \in X$  and  $y = \arg \max_{x \in X} f(x)$  or  $\arg \min_{x \in X} f(x)$ , depending on  $I$ .

**Definition 1.2.2 (Efficient).** We say that an **algorithm**  $A$  is *efficient* if it runs in **polynomial time**.

**Remark (Runtime parametrization).** The *runtime* of an **algorithm**  $A$  should be parametrized by the size of input  $I$ . Formally, given input  $I$  represented in  $s$  bits, runtime of  $A$  on  $I$  should be  $\text{poly}(s)$  for  $A$  to be **efficient**.

**Note.** In most cases, there are 1 or 2 parameters that essentially define the size of input.

**Example (Graph).** A natural representation of a graph with  $n$  vertices and  $m$  edges are

- (a) Adjacency matrix:  $n^2$  numbers.
- (b) Adjacency list:  $O(m + n)$  numbers.

**Example (Set system).** A **set system** with  $n$  elements and  $m$  sets has a natural representation which uses  $O(nm)$  numbers.

**Example.** If an input  $I$  can be represented by  $s$  bits, then the runtime of an **algorithm** can be  $O(s \log s)$ ,  $O(s^2)$ , or  $O(s^{100})$ , which are considered as **efficient**. On the other hand, something like  $2^s$  or  $s!$  are not.

Hence, our goal is to get  $\text{poly}((n, m))$ -time **algorithm**!

## 1.3 Approximation Algorithms

We first note that many interesting **combinatorial optimization problems** are NP-hard, hence it's impossible to find optimum in polynomial time unless P is NP. This suggests one problem: *How well can we do in polynomial time?*

In normal cases, we may assume that objective function value is always positive, i.e.,  $f: X \rightarrow \mathbb{R}^+ \cup \{0\}$ . Then, we have the following definition which characterizes the *slackness*.

**Definition 1.3.1 (Approximation algorithm).** Given an **algorithm**  $A$ , we say  $A$  is an  $\alpha$ -*approximation algorithm* for a **problem**  $P$  if for every input  $I$  of  $P$ ,

- Min:  $f(A(I)) \leq \alpha \cdot \text{OPT}(I)$  for  $\alpha \geq 1$
- Max:  $f(A(I)) \geq \alpha \cdot \text{OPT}(I)$  for  $\alpha \leq 1$

where we define  $\text{OPT}(I)$  as  $\max_{x \in X} f(x)$  for maximization,  $\min_{x \in X} f(x)$  if minimization.

We see that  $\alpha$  characterizes the slackness allowed for our **algorithm**  $A$ . Now, we're ready to look at some interesting **problems**. Broadly, there are around 10 classes of them which are actively studied:

- We'll see cover, clustering, network design, and constraint satisfaction problems.
- We'll not see: graph cuts, Packing, Scheduling, String, etc.

The above list is growing! For example, applications of continuous optimization in **combinatorial optimization** is getting attention recently. Also, there are around 8 techniques developed, e.g., greedy, local search, LP rounding, SDP rounding, primal-dual, cuts and metrics, etc.

## 1.4 Hardness

For most problems we saw, we can even say that getting an  $\alpha$ -approximation is NP-hard for some  $\alpha > 1$ . This bound is sometimes tight, but not always, and we'll focus on this part in the second half of this course.

# Chapter 2

## Covering

### 2.1 Set Cover

Before we jump into any problem formulations, we define a fundamental object in combinatorial optimization, the [set system](#).

**Definition 2.1.1 (Set system).** Given a ground set  $\Omega$  (often called *universe*), the *set system* is an order pair  $(\Omega, \mathcal{S})$  where  $\mathcal{S}$  is a collection of subsets of  $\Omega$ .

**Note.** For a [set system](#)  $(\Omega, \mathcal{S})$ , we often let  $m := |\mathcal{S}|$  and  $n := |\Omega|$ .

**Definition 2.1.2 (Degree).** Given a [set system](#)  $(\Omega, \mathcal{S})$ , the *degree* of  $x \in \Omega$ ,  $\deg(x)$ , is defined as

$$\deg(x) := |\{S \in \mathcal{S} \mid x \in S\}|.$$

**Remark (Bipartite representation).** Naturally, for a [set system](#), we have a bipartite representation.



Figure 2.1: Bipartite representation of a [set system](#).

Denote  $d := \max_{e \in U} \deg(e) \leq m$  and  $k := \max_{i \in [m]} |S_i| \leq n$ , which is just the maximum vertex degree on two sides of the bipartite graph representation of this [set system](#).

Finally, we have the following.

**Definition 2.1.3 (Covering).** A *covering*  $\mathcal{S}' \subseteq \mathcal{S}$  of  $(\Omega, \mathcal{S})$  is a (sub)collection of subsets such that  $\bigcup_{S \in \mathcal{S}'} S = \Omega$ .

Let's first consider the classical problem called [set cover](#).

**Problem 2.1.1 (Set cover).** Given a finite [set system](#)  $(U, \mathcal{S})$  where  $\mathcal{S} := \{S_i \subseteq U\}_{i=1}^m$  along with a weight function  $w: \mathcal{S} \rightarrow \mathbb{R}^+$ , find a [covering](#)  $\mathcal{S}'$  while minimizing  $\sum_{S \in \mathcal{S}'} w(S)$ .

Assuming there always exists at least one [covering](#), we can in fact get two types of non-comparable approximation ratio in terms of  $k$  and  $d$ . Specifically, we get  $\log k$  and  $d$ -approximation ratio via either greedy, LP rounding or dual-methods.

## 2.2 Greedy Method

We first see the algorithm when  $w(S) = 1$  for all  $S \in \mathcal{S}$ .

---

### Algorithm 1: [Set cover](#) – Greedy

---

**Data:** A [set system](#)  $(U, \mathcal{S})$   
**Result:** A [covering](#)  $\mathcal{S}'$

```

1  $\mathcal{S}' \leftarrow \emptyset, i \leftarrow 0$ 
2 while  $U \neq \emptyset$  do                                     //  $O(n)$ 
3   Choose  $S_i$  with maximum  $|U \cap S_i|$                      //  $O(mn)$ 
4   for  $e \in U \cap S_i$  do
5      $y_e \leftarrow w(S_i) / |U \cap S_i|$                      // Average costs
6    $\mathcal{S}' \leftarrow \mathcal{S}' \cup \{S_i\}$ 
7    $U \leftarrow U \setminus S_i$ 
8    $i \leftarrow i + 1$ 
9 return  $\mathcal{S}'$ 

```

---

We focus on the case that  $w(S) = 1$  for all  $S$ .

**Remark.** It's clear that [Algorithm 1](#) is a polynomial time algorithm, also, the output  $\mathcal{S}'$  is always a valid [covering](#).

**Theorem 2.2.1.** [Algorithm 1](#) is an  $H_k$ -approximation<sup>a</sup> algorithm.

<sup>a</sup> $H_k$  is the so-called *harmonic number*, which is defined as  $\sum_{i=1}^k 1/i \leq \ln k + 1$ .

**Proof.** Denote the OPT as  $\mathcal{S}^* := \{S_1^*, \dots, S_\ell^*\}$ , and first note that the average cost  $y_e$  essentially maintains  $\sum_{e \in U} y_e = |\mathcal{S}'|$ , hence we just need to bound  $y_e$  w.r.t.  $\mathcal{S}^*$ . To do this, for any  $S^* \in \mathcal{S}^*$ , say  $S_1^* = \{e_1, \dots, e_k\}$  where we number  $e_i$  in terms of the order of which being deleted, i.e.,  $e_1$  is deleted first from  $U$  ([line 7](#)), etc.

**Note.**  $S_1^*$  can have less than  $k$  element, but in that case similar argument will follow. Also, if some elements are deleted at the same time, we just order them arbitrarily.

Then, we have the following claim.

**Claim.** For all  $e_i$ ,  $y_{e_i} \leq \frac{1}{k-i+1}$ .

**Proof.** Consider the iteration when  $e_i$  was picked by  $\mathcal{S}'$ , i.e.,  $|U \cap \mathcal{S}'| \geq |U \cap S_1^*| \geq k - i + 1$ , then by definition ([line 7](#)) we have  $y_{e_i} = \frac{1}{|U \cap \mathcal{S}'|} \leq \frac{1}{|U \cap S_1^*|} \leq \frac{1}{k-i+1}$ . ⊗

We immediately see that whenever the optimal solution pays 1 (for choosing  $S_1^*$  for instance), [Algorithm 1](#) pays at most  $H_k$  since  $\sum_{e_i \in S_1^*} y_{e_i} \leq \sum_{i=1}^k \frac{1}{k-i+1} = H_k$ , or more formally,

$$|\mathcal{S}'| = \sum_{e \in U} y_e \leq \sum_{S_i^* \in \mathcal{S}^*} \underbrace{\sum_{e \in S_i^*} y_e}_{\leq H_k} \leq \ell \cdot H_k = H_k \cdot |\text{OPT}|,$$

which finishes the proof. ■

In all, observe that  $H_k \leq \ln k + 1$ , we see that [Algorithm 1](#) is a  $(\ln k)$ -approximation algorithm. Also, the weighted version can be easily derived by replacing 1 with the corresponding weight.



## Lecture 2: Linear Programming with Set Covers

### 2.3 Linear Programming Rounding

31 Aug. 10:30

To get a  $d$ -approximation algorithm, instead of seeing the greedy algorithm, we first see the LP<sup>1</sup> dual method, which turns out to be exactly the same as the greedy algorithm.

**As previously seen.** Both linear programming and convex programming can be solved in polynomial time.

Notice that it's more natural to define **set cover** in terms of ILP (integer LP). Define our integer variables  $\{x_i\}_{i \in [n]}$  such that

$$x_i = \begin{cases} 1, & \text{if } S_i \in \mathcal{S}'; \\ 0, & \text{otherwise.} \end{cases}$$

In this way, we have the following ILP formulation for **set cover** as

$$\begin{aligned} \min \quad & \sum_i w_i \cdot x_i \\ & \sum_{S_i \ni e} x_i \geq 1 & \forall e \in U \\ \text{(IP)} \quad & x_i \in \{0, 1\} & \forall i. \end{aligned}$$

But we know that this is a NP-hard problem, so we relax it to be

$$\begin{aligned} \min \quad & \sum_i w_i \cdot x_i \\ & \sum_{S_i \ni e} x_i \geq 1 & \forall e \in U \\ \text{(LP)} \quad & x_i \geq 0 & \forall i. \end{aligned}$$

Write it in a more compact form, we have

$$\begin{aligned} \min \quad & \langle w, x \rangle \\ & Ax \geq \mathbb{1} \\ & x \geq 0 \end{aligned}$$

where  $A \in \mathbb{R}^{n \times m}$  such that

$$A_{ij} = \begin{cases} 1, & \text{if } e_i \in S_j; \\ 0, & \text{otherwise.} \end{cases}$$

**Note.** Note when we do relaxation, we want  $x \in \text{fes}(\text{IP}) \Rightarrow x \in \text{fea}(\text{LP})$ , i.e.,  $\text{fes}(\text{LP}) \supseteq \text{fes}(\text{IP})$ . Note that in this case, for a minimization problem, we have

$$f(x) = \text{OPT}_{\text{LP}} \leq \text{OPT}_{\text{IP}}.$$

In this case, we see that the most natural way to get an integer solution from the fractional solution obtained from the relaxed LP is to **round**  $x$  to integral solution. This leads to the following **algorithm**.

---

**Algorithm 2:** **Set cover** – LP Rounding

---

**Data:** A **set system**  $(U, \mathcal{S})$

**Result:** A **covering**  $\mathcal{S}'$

```

1  $x \leftarrow \text{solve}(\text{LP})$  // Solve the relaxed (LP)
2  $\mathcal{S}' \leftarrow \{S_i : x_i \geq 1/d\}$ 
3 return  $\mathcal{S}'$ 
```

---

We now prove the correctness and **Algorithm 2**'s approximation ratio.

<sup>1</sup>See **MATH561** for a complete reference.

**Lemma 2.3.1.**  $\mathcal{S}'$  is a covering.

**Proof.** Fix  $e \in U$ , let  $S_1, \dots, S_d$  be the sets containing  $e$ . We see that

$$\sum_{i=1}^d x_i \geq 1 \Rightarrow \exists j \in [d] \text{ s.t. } x_j \geq \frac{1}{d} \Rightarrow S_j \in \mathcal{S}'.$$

■

**Theorem 2.3.1.** Algorithm 2 is  $d$ -approximation algorithm.

**Proof.** By comparing  $w(\mathcal{S}')$  and  $\text{OPT}_{\text{LP}} = \sum_{i=1}^m x_i w_i$ , we see that

$$\text{OPT} \leq \sum_{S_i \in \mathcal{S}'} w_i \leq d \sum_{S_i \in \mathcal{S}'} w_i x_i \leq d \cdot \text{OPT}_{\text{LP}} \leq d \cdot \text{OPT},$$

which implies  $\text{OPT} / d \leq \text{OPT}_{\text{LP}} \leq \text{OPT}$ .

**Note.** Note that  $\text{OPT}$  is assumed to be  $\text{OPT}_{\text{IP}}$ , i.e., the optimum of the original IP formulation of Problem 2.1.1.

■

**Definition 2.3.1 (Integrality gap).** Given an integer programming, the *integrality gap* between  $\text{OPT}$  and  $\text{OPT}_{\text{LP}}$  of its LP relaxation is defined as

$$\sup_{\text{input } I} \frac{\text{OPT}(I)}{\text{OPT}_{\text{LP}}(I)}.$$

**Remark.** We see that the integrality gap of Algorithm 2 is  $d$  from Theorem 2.3.1.

### 2.3.1 Randomized Linear Programming Rounding

And indeed, we can use a more natural way to do the rounding, i.e., respect to the  $x_i$  value.

**Intuition.** If  $x_i$  is close to 1, it's reasonable to include it, vice versa.

We see that algorithm first.

---

**Algorithm 3:** Set cover – Randomized LP Rounding

---

**Data:** A set system  $(U, \mathcal{S})$

**Result:** A (possible) covering  $\mathcal{S}'$

```

1  $x \leftarrow \text{solve}(\text{LP})$  // Solve the relaxed (LP)
2  $\mathcal{S} \leftarrow \emptyset$ 
3 for  $i = 1, \dots, m$  do
4    $\lfloor$  add  $S_i$  to  $\mathcal{S}'$  w.p.  $x_i$  // independently
5 return  $\mathcal{S}'$ 
```

---

Now, the question is, how is this  $\mathcal{S}'$ 's quality? In other words, fix  $e \in U$ , what's  $\Pr(e \text{ is covered})$ ?

**Lemma 2.3.2.**  $\Pr(e \text{ is covered}) \geq 1 - 1/e \approx 0.63$ .

**Proof.** We bound  $\Pr(\overline{e \text{ is covered}})$  instead. Say  $S_1, \dots, S_d$  are the sets containing  $e$ , then we see

that

$$\Pr(\overline{e \text{ is covered}}) = \prod_{i=1}^d (1 - x_i) \leq \prod_{i=1}^d e^{-x_i} = e^{-\overbrace{(x_1 + \dots + x_d)}^{\geq 1}} \leq e^{-1}.$$

**Note.** For every  $x$ , we have  $1 + x \leq e^x$ , and this approximation is close when  $|x|$  is small. ■

A standard way to boost the correctness of a randomized algorithm is to run it multiple time, which leads to the following.

---

**Algorithm 4:** Set cover – Multi-time Randomized LP Rounding

---

**Data:** A set system  $(U, \mathcal{S})$ ,  $\alpha$

**Result:** A (possible) covering  $\mathcal{S}'$

```

1  $x \leftarrow \text{solve}(\text{LP})$  // Solve the relaxed (LP)
2  $\mathcal{S} \leftarrow \emptyset$ 
3 for  $t = 1, \dots, \alpha$  do // independently
4   for  $i = 1, \dots, m$  do
5      $\sqcup$  add  $S_i$  to  $\mathcal{S}'$  w.p.  $x_i$  // independently
6 return  $\mathcal{S}'$ 
```

---

**Lemma 2.3.3.** With  $\alpha = 2 \ln n$ ,  $\mathcal{S}'$  returned from Algorithm 4 is a covering w.p. at least  $1 - \frac{1}{n}$ .

**Proof.** We have  $\Pr(e \text{ is not covered}) \leq e^{-\alpha}$  from independence of each run. Let  $\alpha = 2 \ln n$ , then  $\Pr(e \text{ is not covered}) \leq e^{-\alpha} = 1/n^2$ . By union bound,

$$\Pr(\text{some elements is not covered}) \leq \sum_{e \in U} \Pr(e \text{ not covered}) \leq n \cdot \frac{1}{n^2} = \frac{1}{n}.$$

This implies w.p. at least  $1 - 1/n$ ,  $\mathcal{S}'$  is a covering. ■

In other words, with  $\alpha = 2 \ln n$ , Algorithm 4 is correct with probability at least  $1 - 1/n$ .

**Lemma 2.3.4.** With  $\alpha = 2 \ln n$ ,  $\mathcal{S}'$  returned from Algorithm 4 has an approximation ratio  $4 \ln n$  w.p. at least  $\frac{1}{2}$ .<sup>a</sup>

<sup>a</sup>Note that  $\mathcal{S}'$  is not necessary a covering.

**Proof.** Since  $\mathbb{E}[w(\mathcal{S}')] \leq \alpha \sum_i w_i x_i = \alpha \text{OPT}_{\text{LP}}$ , we have  $\Pr(w(\mathcal{S}') \geq 2 \cdot \alpha \text{OPT}_{\text{LP}}) \leq 1/2$  from Markov inequality. We see that w.p.  $\geq 1/2$ ,  $w(\mathcal{S}') \leq 2 \cdot 2 \ln n \cdot \text{OPT}_{\text{LP}} \leq 4 \ln n \text{OPT}$ . ■

**Theorem 2.3.2.** By running Algorithm 4 many times, we get a  $(4 \ln n)$ -approximation algorithm with high probability.<sup>a</sup>

<sup>a</sup>Note that we still need to choose  $\mathcal{S}'$ .

**Proof.** Together with Lemma 2.3.3 and Lemma 2.3.4 and using the union bound, the probability of  $\mathcal{S}'$  not being a covering or with weight higher than  $4 \ln n \text{OPT}$  is at most  $\frac{1}{n} + \frac{1}{2}$ , which is less than 1. Hence, by running Algorithm 4 many times (independently), the failing possibility is exponential small. ■

**Note.** With Theorem 2.3.2, we still need to find a valid covering with the lowest cost, where a valid covering with low enough weight is guaranteed to exist with high probability. Note that this is still a polynomial time algorithm since we know that checking  $\mathcal{S}'$  is a covering is just linear.

**Remark.** Indeed, with some smarter algorithm modified from [Algorithm 4](#), we can get an  $H_k$  approximation ratio.

## Lecture 3: Covering-Packing Duality and Primal-Dual Method

### 2.4 Covering-Packing Duality

7 Sep. 10:30

We first define some useful notions.

**Definition 2.4.1 (Strongly independent).** Given a set system  $(U, \mathcal{S})$ , we say  $C \subseteq U$  is *strongly independent* if there does not exist  $S \in \mathcal{S}$  such that  $|C \cap S| \geq 2$ .

**Remark.** Then for any [strongly independent](#) set  $C \subseteq U$ , we know that  $\text{OPT}_{\text{SC}} \geq |C|$ .<sup>a</sup>

<sup>a</sup>SC denotes [set cover](#).

Now, we're trying to find the **strongest witness** of [strongly independent set](#), which suggests we define the following problem.

**Problem 2.4.1 (Maximum strongly independent set).** Given a set system  $(U, \mathcal{S})$ , we want to find the largest [strongly independent set](#).

**Remark.** For any set system, we have  $\text{OPT}_{\text{SIS}} \leq \text{OPT}_{\text{SC}}$ .<sup>a</sup>

<sup>a</sup>SIS denotes [maximum strongly independent set](#).

**As previously seen (LP dual).** Recall how we get the dual of a given LP:

$$\begin{array}{ll} \min & c^\top x \\ & Ax \geq b \\ (P) & x \geq 0 \end{array} \qquad \begin{array}{ll} \max & y^\top b \\ & y^\top A \leq c^\top \\ (D) & y \geq 0. \end{array}$$

Also, recall the weak duality ( $\text{OPT}_P \geq \text{OPT}_D$ ) and strong duality ( $\text{OPT}_P = \text{OPT}_D$ ).

**Definition 2.4.2 (Covering LP).** A primal LP with  $A, b, c \geq 0$  is called a *covering LP*.

**Definition 2.4.3 (Packing LP).** A dual LP with  $A, b, c \geq 0$  is called a *packing LP*.

We now give another LP formulation for the unweighted [set cover](#). Given  $\mathcal{S} = \{S_1, \dots, S_m\}$ ,  $U = \{e_1, \dots, e_n\}$  and define  $A \in \mathbb{R}^{n \times m}$  such that

$$A_{ij} = \begin{cases} 1, & \text{if } e_i \in S_j; \\ 0, & \text{otherwise.} \end{cases}$$

Then our LP is defined as

$$\begin{array}{ll} \min & \sum_{j=1}^m x_j \\ & Ax \geq \mathbf{1} \\ (P) & x \geq 0 \end{array} \qquad \begin{array}{ll} \max & \sum_{i=1}^n y_i \\ & y^\top A \leq \mathbf{1} \\ (D) & y \geq 0. \end{array}$$

We see that if we restrict  $y_i \in \{0, 1\}$ , we see that the dual (D) is just [Problem 2.4.1](#). This can be

seen via writing the constraint explicitly:

$$\sum_{i=1}^n A_{ij} y_i \leq 1 \Leftrightarrow \sum_{i: e_i \in S_j} y_i \leq 1 \text{ for } j \in [m].$$

And indeed, if we look at the weighted version, we have  $\sum_{i: e_i \in S_j} y_i \leq w(S_j)$ .

Now, recall the claim in Theorem 2.2.1, i.e.,  $y_{e_i} \leq \frac{w(S_j)}{k-i+1}$ . We see that the  $y_{e_i}$  are just the dual variables in our setup. Additionally, with the observation that we can do this for any set  $S = \{e_1, \dots, e_k\} \in \mathcal{S}$ , we have the following lemma.

**Lemma 2.4.1.** The variable  $y' := y/H_k$  is dual-feasible, i.e., it's feasible for  $(D)$ .

**Proof.** We see that  $y_{e_i} \geq 0$  (and hence  $y_i$ ) trivially, so we only need to show that

$$\sum_{i=1}^n A_{ij} y' = \sum_{i=1}^n A_{ij} \frac{y_{e_i}}{H_k} \leq w(S_j)$$

for  $j \in [m]$ . But this is trivial by plugging in  $y_{e_i} \leq \frac{w(S_j)}{k-i+1}$  as shown in Theorem 2.2.1, hence

$$\sum_{i=1}^n A_{ij} \frac{y_{e_i}}{H_k} \leq \frac{1}{H_k} \sum_{i=1}^n A_{ij} \frac{w(S_j)}{k-i+1} \leq \frac{1}{H_k} \sum_{i=1}^k \frac{w(S_j)}{k-i+1} = w(S_j),$$

and we're done.<sup>a</sup> ■

<sup>a</sup>Note that in the above derivation,  $i$  is kind of overloading, i.e.,  $e_i$  corresponding to only some  $i$  (confusing, but it's how it is...).

With Lemma 2.4.1, we simply run Algorithm 1 while maintaining  $y_e$  for every  $e$ , and we're done.

**Theorem 2.4.1.** Algorithm 1 is an  $H_k$ -approximation algorithm in the view of its dual.

**Proof.** Same as Theorem 2.2.1, but now we have different interpretation. Specifically, if  $y' = y/H_k$  is dual-feasible, we know that the corresponding objective value of  $y'$  is at most  $\text{OPT}_{\text{LP}_D} = \text{OPT}_{\text{LP}_P}$ , which is at most  $\text{OPT}_{\text{SC}}$  further. Now, since we're dealing with LP, everything is linear includes the objective value, i.e.,  $y$  is at most  $H_k \cdot \text{OPT}_{\text{SC}}$ . ■

**Remark (Dual fitting).** The above method is called *dual fitting*, which is universal as one can easily see. The way to do this is the following.

1. Given an algorithm, distribute the algorithm to  $\{y_i\}$ .
2. Prove that  $y/\alpha$  is dual-feasible.
3. This shows the algorithm is  $\alpha$ -approximation algorithm.

## 2.5 Primal-Dual Method

We first see the general description of the so-called *primal-dual method*.

1. Maintain  $x$  (primal solution) and  $y$  (dual solution) where  $x$  is integral and infeasible, while  $y$  is fractional and feasible. Start from  $x = y = 0$ .
2. **Somehow** increase  $y$  until some dual constraints get tight.
3. **Choose** primal variables correspond to tight dual constraints, and update input accordingly.

**Remark.** We're using dual variables to get a certificate of the lower bound of the optimal problem we're solving.

In terms of [set cover](#), we have the following.

---

**Algorithm 5:** [Set cover](#) – Primal-Dual
 

---

**Data:** A [set system](#)  $(U, \mathcal{S})$

**Result:** A [covering](#)  $\mathcal{S}'$

```

1  $\mathcal{S}' \leftarrow \emptyset, y \leftarrow 0$ 
2 while  $U \neq \emptyset$  do
3   Choose any  $e \in U$ 
4   Raise  $y_e$  until some constraints get tight
5    $\mathcal{S}' \leftarrow \mathcal{S}' \cup \{\text{sets corresponding to tight dual constraints}\}$ 
6   Update  $U$                                      // Remove newly covered element in  $U$ 
7 return  $\mathcal{S}'$ 
  
```

---

**Remark.** [Algorithm 5](#) is correct and can be implemented efficiently.

**Theorem 2.5.1.** [Algorithm 5](#) is a  $d$ -approximation algorithm.

**Proof.** Firstly,  $y$  is feasible. And we see that

$$w(\mathcal{S}') = \sum_{S \in \mathcal{S}'} w(S) = \sum_{S \in \mathcal{S}'} \sum_{e \in S} y_e \leq d \cdot \sum_{e \in U} y_e \leq d \cdot \text{OPT}_{\text{LP}_D} = d \cdot \text{OPT}_{\text{LP}_P} \leq d \cdot \text{OPT}_{\text{SC}}.$$

■

## Lecture 4: Feedback Vertex Set

### 2.6 Feedback Vertex Set

12 Sep. 10:30

Following the discussion on primal-dual method, we see another [covering](#) problem.

#### 2.6.1 Introduction

We consider the following problem.

**Problem 2.6.1** ([Feedback vertex set](#)). Given a graph  $\mathcal{G} = (\mathcal{V}, \mathcal{E})$  and a weight function  $c: \mathcal{V} \rightarrow \mathbb{R}^+$ , we want to find  $F \subseteq \mathcal{V}$  with  $\min c(F)$  such that  $\mathcal{G}[\mathcal{V} \setminus F]$  has no cycle.<sup>a</sup>

<sup>a</sup>This is equivalent as saying that  $\mathcal{G}[\mathcal{V} \setminus F]$  is a forest.

**Note** ([Feedback](#)). The name *feedback* comes from the fact that if there's a cycle in  $\mathcal{G}$ , then it kind of creates *feedback*.

**Note** ([Edge version](#)). The *edge version* of [Problem 2.6.1](#) can be solved by finding  $T \subseteq \mathcal{E}$  be the maximum weight forest,<sup>a</sup> and let  $F := \mathcal{E} \setminus T$ .

<sup>a</sup>This can be found exactly in polynomial time.

**Notation.** In this lecture, when talking about cycle, we're referring to the vertices in which. But the meaning can vary from context to context.

**Remark.** This is a *special case* of [Problem 2.1.1](#).

**Proof.** Let  $\mathcal{C} := \{\text{set of all (simple) cycles}\}$  and consider [Problem 2.1.1](#) on the set system  $(\mathcal{C}, \mathcal{V})$ , i.e., we want to find  $F \subseteq \mathcal{V}$  such that  $\forall C \in \mathcal{C}, |F \cap C| \geq 1$ .  $\circledast$

**Note.** The naive algorithm by directly applying methods discussed for [Problem 2.1.1](#), we see that since  $\min(\log k, d) = \Omega(n)$  for  $k$  being the maximum set size (which is  $2^{\Omega(n)}$ ) and  $d = n$ , the approximation ratio we can get is  $\Omega(n)$ , which depends on the size of the input.

Now, the goal in this section is to show the following.

**Theorem 2.6.1.** There exists a 4-approximation algorithm for [Problem 2.6.1](#).

**Remark.** Actually, there exists a 2-approximation algorithm.

We also have a hardness of [Problem 2.6.1](#).

**Theorem 2.6.2.** Achieving  $(2 - \epsilon)$ -approximation algorithm if NP-hard for all  $\epsilon > 0$  assuming the unique games conjecture.

**Proof.** See Homework 1. ■

## 2.6.2 Cycle Covering LP

The most natural LP which models [Problem 2.6.1](#) is the so-called *cycle covering LP*, which can be defined as

$$\begin{aligned} \min \quad & \sum_{v \in \mathcal{V}} c(v)x_v \\ & \sum_{v \in C} x_v \geq 1 \quad \forall \text{ cycle } C \in \mathcal{C} \\ & x \geq 0, \end{aligned}$$

with the variables being  $\{x_v\}_{v \in \mathcal{V}}$  such that  $x_v = \mathbb{1}_{v \in F}$ .

**Remark.** We see that this cycle covering LP has  $2^{\Omega(n)}$  constraints. But we can actually solve this and get an  $O(\log n)$ -approximation ratio by smartly rounding the solution.<sup>a</sup> And we can show that this approximation ratio is optimal in terms of this particular LP.

<sup>a</sup>See homework 1.

## 2.6.3 Density LP

A more sophisticated LP is the so-called *density LP*, defined as

$$\begin{aligned} \min \quad & \sum_{v \in \mathcal{V}} c(v)x_v \\ & \sum_{v \in S} x_v (d_v^S - 1) \geq |E(S)| - |S| + 1 \quad \forall S \subseteq \mathcal{V} \\ & x \geq 0 \end{aligned}$$

with the variables being  $\{x_v\}_{v \in \mathcal{V}}$ .

**Notation.** The  $E(S)$  denotes the edge set in the induced graph  $\mathcal{G}[S] = (S, E(S))$ , while  $d_v^S$  denotes the degree of  $v$  in  $\mathcal{G}[S]$ .

**Intuition.** The constraint is equivalent as saying that for every induced graph,  $\#e \leq \#v - 1$ , i.e., we require it to be a forest. Explicitly,  $S \subseteq \mathcal{V}$ ,

$$|E(S)| - \sum_{v \in S} x_v d_v^S \leq |S| - \sum_{v \in S} x_v - 1.$$

Note that in the constraint, the right-hand side is just a lower-bound of  $\#e$ .

We see that the above LP is not exactly a **covering LP** since the coefficients can be negative if a set  $S$  is not **irreducible**.

**Definition 2.6.1 (Irreducible).** The set  $S \subseteq \mathcal{V}$  is *irreducible* if for all  $v \in S$ ,  $v$  belongs to some cycles in  $G[S]$ .

Now, it's clear that by looking at  $\mathcal{S} = \{S \subseteq \mathcal{V} \mid S \text{ is irreducible}\}$ , we have a **covering LP** defined as

$$\begin{aligned} \min \quad & \sum_{v \in \mathcal{V}} c(v)x_v \\ & \sum_{v \in S} x_v(d_v^S - 1) \geq |E(S)| - |S| + 1 =: b_S \quad \forall S \in \mathcal{S} \\ & x \geq 0. \end{aligned}$$

We first see why this LP models **Problem 2.6.1**.

**Lemma 2.6.1.** The integer version of density LP (denote as IP) is equivalent to **Problem 2.6.1**.

**Proof.** If  $x$  is feasible for **Problem 2.6.1**, then  $x$  is feasible for the IP. On the other hand, if  $x$  is feasible for IP, then for every cycle  $C \in \mathcal{C}$ ,  $x$  deletes at least 1 vertex from  $C$ . ■

## 2.6.4 Primal-Dual Method

Now we're ready to solve this LP via primal-dual method. Denote the dual variables as  $\{y_S\}_{S \in \mathcal{S}}$ , then the dual is

$$\begin{aligned} \max \quad & \sum_{S \in \mathcal{S}} y_S b_S \\ & \sum_{S \ni v} (d_v^S - 1)y_S \leq c(v) \quad \forall v \in \mathcal{V} \\ & y \geq 0. \end{aligned}$$

**Note.** For the density LP and its dual, the constraint is still exponentially many, and no one knows how to solve this. But the power of primal-dual method is that we don't really solve this, rather, we just maintain two sets of solutions for both primal and dual. Moreover, we can maintain the primal solution in integral, while the dual solution in fractional.

We now have the following algorithm.



**Algorithm 6:** Feedback vertex set – Primal-Dual

---

**Data:** A graph  $\mathcal{G} = (\mathcal{V}, \mathcal{E})$   
**Result:** A minimal feedback vertex set  $F'$

```

1  $S \leftarrow \mathcal{V}, c' = c, y \leftarrow 0$  //  $c' \in \mathbb{R}^n$  keeps track of slackness of  $c$ 
2
3 while  $S \neq \emptyset$  do
4    $S \leftarrow \text{reduce}(S)$  // Compute  $\{v \in S: v \text{ belongs to some cycles in } \mathcal{G}[S]\}$ 
5    $(\alpha, v) \leftarrow \min_{v \in S} c'(v)/(d_v^S - 1)^a$  //  $y_S$  gets tight by increasing unit weight
6    $y_S \leftarrow \alpha$ 
7    $c'(v) \leftarrow c'(v) - \alpha(d_v^S - 1)$ 
8    $Z \leftarrow \{v \in S: c'(v) = 0\}$ 
9    $F \leftarrow F \cup Z, S \leftarrow S \setminus Z$ 
10
11 // Compute a minimal feedback vertex set
12  $F' \leftarrow F = \{v_1, \dots, v_\ell\}$  //  $v_1$  is deleted first,  $v_\ell$  is deleted last
13 for  $i = \ell, \dots, 1$  do // reversed greedy
14   if  $F' \setminus \{v_i\}$  is a feedback vertex set for  $\mathcal{G}$  then
15      $F' \leftarrow F' \setminus \{v_i\}$ 
16 return  $F'$ 

```

---

<sup>a</sup>Note that we also get the argument  $v$ .

We see that in Algorithm 6, we first use primal-dual method to obtain a feasible feedback vertex set, and then run a reversed greedy algorithm to further ensure we get a good approximation ratio.

**Claim.**  $F$  is a feedback vertex set and  $y$  is dual-feasible.

**Proof.** It should be clear that why  $F$  is a feedback vertex set. As for the reason why  $y$  is dual-feasible, observe that we have one constraint for each  $v$ . After raising  $y_S$  for chosen  $v$  in line 6 and deduce  $c'(v)$  in line 7,  $v$  will get removed so the constraint corresponding to  $v$  will be satisfied throughout. \*

**Remark (Reversed greedy).** The method we turn  $F$  into its minimal is called *reversed greedy*. This just checks that if we remove a vertex  $v$  from  $F'$  while  $F'$  is still feasible, then we just do it. Additionally, we iterate through  $v$  in the **reversed** order w.r.t. how  $v$  is being added into.

We want to compare the primal cost and the dual cost. The primal cost is

$$c(F) = \sum_{v \in F} c(v) = \sum_{v \in F} \sum_{S \ni v} (d_v^S - 1) y_S = \sum_{S \in \mathcal{S}} y_S \sum_{v \in F \cap S} (d_v^S - 1),$$

while the dual cost is  $\sum_{S \in \mathcal{S}} y_S b_S$ .

**Remark.** This is where the primal-dual method is powerful. i.e., by switching the order of summation, if we have some ratio of  $\sum_{v \in F \cap S} (d_v^S - 1)$  and  $b_S$  for every  $S$ , we're done. On caveat is that since  $S$  is changing when running Algorithm 6, so the final solution  $F$  may not be good for this particular  $S$ . We need to guarantee some ratio for this  $F$  **for all**  $S$ .<sup>a</sup>

<sup>a</sup>At least for  $S$  with positive  $y_S$ .

**Lemma 2.6.2.** For all  $S \in \mathcal{S}$ , if  $F$  is *minimal* in  $S$ ,<sup>a</sup> then we have

$$\sum_{v \in S \cap F} (d_v^S - 1) \leq 4 \cdot b_S = 4(|E(S)| - |S| + 1).$$

<sup>a</sup>i.e., in  $\mathcal{G}[S]$ , no  $F' \subsetneq F \cap S$  in feedback vertex set.

**Proof.** Let's first see a simple case.

**Intuition.** If the graph is 3-regular, then we see that the left-hand side is  $\leq 2 \cdot |S|$  by summing over the whole  $S$  instead of  $S \cap F$ , while the right-hand side is  $2 \cdot |S| + 4$  since  $|E(S)| = 1.5 |S|$ .

This shows that in a 3-regular graph, deleting every vertex in  $S$  is actually 4-approximated. And this intuition generalized to general graph with degree greater than 3.

Since we assume  $S$  to be **irreducible**, so we're not interested in degree 0 or 1 vertices (there are no such vertices in an **irreducible**  $S$ ). So the only problematic guy is degree-2 vertex. And the only place a degree-2 vertex can live is in a long path.



Figure 2.2: If there are two  $v \in F$ , by minimality of  $F$ , one of  $v$  will be strictly unnecessary to break this path in a cycle.

**Note.** Observe that we only need to delete at most one vertex in any path, and sometimes this may be loose since we can delete one branch node joining two paths, i.e., deleting 1 nodes for two paths.

Let  $A$  be the set of degree 2 vertices, and  $B$  be the set of vertices with degree larger than 3. Now, consider line segment in the graph. If  $\ell$  is a line segment,

- (a)  $|F \cap \ell| \leq 1$ , i.e., we delete at most one point in  $\ell$ .
- (b) If  $F$  contains one of the endpoints of  $\ell$ , then  $|F \cap \ell| = 0$ .

Since  $F$  is minimal, the left-hand side is

$$|A \cap F| + \sum_{v \in B \cap F} (d_v^S - 1) \leq \sum_{v \in B \setminus F} d_v^S / 2 + \sum_{v \in B \cap F} (d_v^S - 1) \leq \sum_{v \in B} (d_v^S - 1),$$

where the first inequality comes from the fact that if we delete vertices in  $A$ , i.e., in the line segment, then we know we don't delete its end points, and by *distributed* that 1 cost into its two end points, each  $1/2$ .



Figure 2.3: Distribute the cost of  $F$ .

Similarly, in the right-hand side, the crucial term is

$$|E(S)| - |S| = \sum_{v \in S} (d_v^S / 2 - 1) = \sum_{v \in B} (d_v^S / 2 - 1)$$

where the last equality holds since for  $v \in A$ , the summand is just  $2/2 - 1 = 0$ . It's clear that since  $\forall v \in B, d_v^S - 1 \leq 4(d_v^S / 2 - 1)$ , rearranging this inequality gives the result. ■

To show [Theorem 2.6.1](#), it's enough to have a minimal  $F$ , then the result follows from [Lemma 2.6.1](#). Hence, after obtaining  $F$ , [Algorithm 6](#) further convert  $F$  into  $F'$  and try to obtain a minimal version of  $F$ . Clearly,  $F'$  is still a **feedback vertex set**, and the minimality of  $F'$  is guaranteed by the following

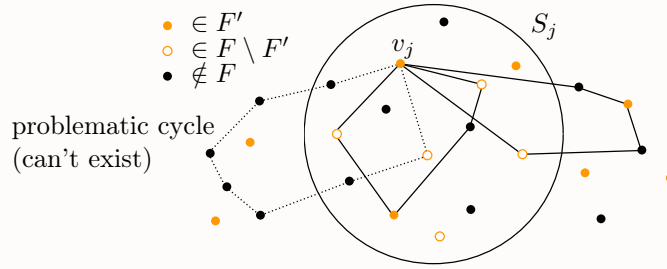
lemma.

**Lemma 2.6.3.**  $F'$  is minimal in every  $S_i$ , where  $S_i$  is the corresponding  $S$  in Algorithm 6 when  $v_i$  is deleted.

**Proof.** Suppose this is not the case. Then there exists  $v_j \in F'$  such that in  $\mathcal{G}[S_i]$ ,  $(F \cap S_i) \setminus \{v_j\}$  is still a **feedback vertex set** in  $\mathcal{G}[S_j]$ . Notice that we only need to consider the case that  $i = j$  since  $v_j \in S_i$  means  $i \geq j$  from how we order them. In this case,  $S_j \subseteq S_i$ , hence to check the minimality of  $F'$  it's enough to just consider the case that  $i = j$ . Hence, we consider  $(F \cap S_j) \setminus \{v_j\}$  instead.

**Note.** Here we only consider  $G[S_j]$ , i.e., we want to say that if  $v_j$  is not minimal in  $G[S_j]$ , then  $v_j$  should really be deleted even w.r.t. the whole graph.

Now, observe the following picture in step  $j$  of line 13 with cycles contained  $v_j$ :



Observe that the middle cycles in  $G[S_j]$  must exist from our assumption of  $(F \cap S_j) \setminus \{v_j\}$  being still a **feedback vertex set**, i.e., if a cycle exists in  $G[S_j]$ , then it must contain another nodes other than  $v_j$  that's also in  $F'$ . But we see that when we consider cycles outside  $G[S_j]$ , we have the following.

**Claim.** No vertices outside  $S_j$  which is also in  $F \setminus F'$  at step  $j$  of line 13

**Proof.** Since  $S_j$  is growing, i.e., for  $i \leq j$ ,  $S_i \subseteq S_j$ , and we just can't delete something we haven't considered.  $\otimes$

**Claim.** There are no cycles  $C \ni v_j$  such that  $C \setminus S_j$  is disjoint from  $F$ .

**Proof.** Observe that there are only two ways for a vertex being deleted from the graph, either  $v \in Z$ , i.e., its dual constraint is tight, or  $v \in S$  is deleted since it prevent  $S$  being **irreducible**. Only the latter case will make  $v \notin F$ , we see that there's no way such a cycle  $C$  exists with all vertices outside  $S_j$  are preventing  $s$  being **irreducible**, since this cycle  $C$  itself is a cycle...  $\otimes$

This implies  $F' \setminus \{v_j\}$  is still a **feedback vertex set** in  $\mathcal{G}$  when  $i = j$  in Algorithm 6 since such a problematic cycle can't exist,<sup>a</sup> which contradicts with the minimality of  $F'$ . ■

<sup>a</sup>Explicitly, if this exists, then delete  $v_j$  will make  $F'$  fail to intersect such a cycle.

Finally, we see that we can prove Theorem 2.6.1.

**Proof of Theorem 2.6.1.** Firstly, Algorithm 6 gives a 4-approximation of the density IP guaranteed by Lemma 2.6.2 and Lemma 2.6.3. Finally, from Lemma 2.6.1, we see that Problem 2.6.1 and the density IP is equivalent, proving the theorem. ■

# Chapter 3

## Clustering

### Lecture 5: Facility Location

#### 3.1 Clustering

14 Sep. 10:30

The problem we're interested in is called the **clustering** problem.

**Problem 3.1.1 (Clustering).** Given  $n$  objects, partition them into  $k$  groups such that

- *Similar* objects are in same group
- *Different* objects are in different group.

**Note.** We see that **Problem 3.1.1** is vague in terms of the definition, which is because this is more like a class of problems. We'll see different notions of *similar* and *different* later when we consider more explicit problems.

In particular, the notion of **metric** is useful.

**Definition 3.1.1 (Metric).** Given a set  $X$ , a function  $d: X \times X \rightarrow \mathbb{R}^+ \cup \{0\}$  is called a *metric* if

- (a)  $d(\cdot, \cdot) \geq 0$  and  $d(i, j) = 0$  if and only if  $i = j$ .<sup>a</sup>
- (b)  $d(i, j) = d(j, i)$  for all  $i, j \in X$ .
- (c)  $d(i, j) + d(j, k) \geq d(i, k)$  for all  $i, j, k \in X$ .

<sup>a</sup>We didn't mention this in lectures, but in math community this should also be included.

**Remark (Metric space).** Though we didn't formally introduce, but the pair  $(X, d)$  of  $X$  and a **metric**  $d$  on  $X$  is sometimes called a *metric space*.

#### 3.2 Facility Location

Let's first look at the problem.

**Problem 3.2.1 (Facility location).** Given a **metric space**  $(X, d)$  and  $P, Q \subseteq X$ ,  $f \in \mathbb{R}^+$  where  $P$  is the set of clients,  $Q$  is the set of (possible) facilities, we want to open  $O \subseteq Q$  such that it minimizes

$$\sum_{i \in P} \min_{j \in O} d(i, j) + f |O|,$$

where we interpret the first summation as connection cost, the second term as opening cost.

**Example.** Consider the following example.



If  $f = 1$  and we open the black facilities, then the cost is  $2 + 5 = 7$  assuming unit length.

We now write down the LP of [Problem 3.2.1](#). Denote variables  $\{y_j\}_{j \in Q}$  and  $\{x_{ij}\}_{i \in P, j \in Q}$ . Then the LP can be written as

$$\begin{aligned}
 \min \quad & \sum_{ij} d(i, j) x_{ij} + \sum_j y_j \cdot f \\
 & \sum_j x_{ij} \geq 1 & \forall i \in P & (\alpha_i) \\
 & x_{ij} \leq y_j \Leftrightarrow y_j - x_{ij} \geq 0 & \forall i, j & (\beta_{ij}) \\
 (P) \quad & x, y \geq 0.
 \end{aligned}$$

Denote the dual variables as  $\alpha_i$  and  $\beta_{ij}$ , the dual is

$$\begin{aligned}
 \max \quad & \sum_i \alpha_i \\
 & \alpha_i - \beta_{ij} \leq d(i, j) & \forall i, j & (x_{ij}) \\
 & \sum_i \beta_{ij} \leq f & \forall j & (y_i) \\
 (D) \quad & \alpha, \beta \geq 0.
 \end{aligned}$$

**Remark.** If  $(\alpha, \beta)$  is feasible, redefine  $\beta_{ij} := \max(0, \alpha_i - d(i, j))$ , it's still feasible and will not affect the objective value. We see that we can drop  $\beta$  and only look at  $\alpha$ .

We can then define the following useful notion called [cluster](#).

**Definition 3.2.1 (Cluster).** A *cluster*  $C := (j, P')$  is the order pair for  $j \in Q$  and  $P' \subseteq P$ , where the cost  $c(C)$  is calculated by directing all  $i \in P'$  to  $j$ , i.e.,  $c(C) = f + \sum_{i \in P'} d(i, j)$ .

**Notation.** We denote the set of all [clusters](#)  $C$  by  $\mathcal{C}$ .

**Remark (Just set cover!).** We see that [Problem 3.2.1](#) is equivalent to [set cover](#) on  $(P, \mathcal{C})$ .

**Proof.** If we write down the LP for [set cover](#) on  $(P, \mathcal{C})$ , we have

$$\begin{aligned}
 \min \quad & \sum_{C \in \mathcal{C}} c(C) \cdot y_C & \max \quad & \sum_{i \in P} \alpha_i \\
 & \sum_{C \ni i} y_C \geq 1 \quad \forall i \in P & & \sum_{i \in C} \alpha_i \leq c(C) \quad \forall C \in \mathcal{C} \\
 (P) \quad & y \geq 0 & (D) \quad & \alpha \geq 0,
 \end{aligned}$$

which is equivalent to what we have as above. \*

But observe that the number of clusters is  $|Q| \cdot 2^{|P|}$ , hence directly solve either (P) or (D) is not feasible. In this case, we can use the primal-dual method.

### 3.2.1 Primal-Dual Method

Let's first see the primal-dual algorithm on  $(P)$  and  $(D)$  derived above.

---

**Algorithm 7:** Facility location – Primal-Dual
 

---

**Data:** A set of clients  $P \subseteq X$ , a set of (possible) facilities  $Q \subseteq X$ , facility cost  $f$   
**Result:** A set of opened facilities  $O \subseteq Q$

```

1  $S \leftarrow \emptyset, O \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $O$ :open facilities
2
3 while  $S \neq P$  do
4   while True do
5     increase all  $\{\alpha_i\}_{i \in P \setminus S}$  by a unit
6     if some  $j \in Q \setminus O$  s.t.  $\sum_{i \in P} \beta_{ij} = f$  then //  $j$  gets tight (open)
7       break
8     else if some  $i \in P \setminus S$  s.t.  $\alpha_i \geq d(i, j)$  then //  $i$  can connect to  $j \in O$ 
9       break
10     $O \leftarrow \{\text{tight facilities}\}$  // Update  $O$ 
11     $S \leftarrow \{\text{clients connected to } O\}$  // Update  $S$ 
12
13 // Trim down  $O$ 
14  $G = (O, E := \{(j, j') : \exists i \in P \text{ such that } \alpha_i > d(i, j), j, j' \in O\})$ 
15 Compute  $O'$  s.t.  $\forall j \in O$ , either  $j \in O'$  or  $\exists j' \in O'$  s.t.  $(j, j') \in E$  // maximal independent set
16 return  $O'$ 
```

---

**Note.** line 6 and line 8 can happen in the same time.

**Intuition.** We're basically increasing the cost  $i$  willing to pay and stop (in the second while loop) when  $i$  finally connect to  $j$ . Or one can also interpret  $\alpha_i$  as the time  $i$  connects to some facilities  $j$ .

This directly relates to the fact that for all  $i, j$ , if  $i, j$  are connected, then  $d(i, j) \leq \alpha_i$ , which is exactly the spirit of the primal-dual method since we want to argue the upper-bound in terms of  $\alpha$ . But before that, we need to argue that  $\alpha$  is actually feasible in order to make this bound valid.

**Lemma 3.2.1.**  $\alpha$  is dual-feasible in Algorithm 7.

**Proof.** Firstly,  $\alpha$  start from 0 which is feasible. Now, for  $\alpha_i$  violates the constraints  $\sum_{i \in C} \alpha \leq c(C) = f + \sum_{i \in P'} d(i, j)$ , there are two possibilities, but both are handled in Algorithm 7. Specifically, line 6 and line 8:

- In line 6: This corresponds to some  $j$  gets opened, we then need to make sure that no  $\alpha_i$  will pay toward  $j$  for its open cost  $f$ . But this is clear since whoever  $i$  is paying non-zero amounts to  $j$  for its  $f$ ,  $i$  immediately connect to  $j$  and will be clicked out from  $P \setminus S$ , meaning that their dual  $\alpha_i$  will not be increased anymore.
- In line 8: This corresponds to when  $i$  want to connect (willing to pay non-zero amount to) an already opened  $j$ . But we see that whenever  $i$  willing to pay for an already opened  $j$ , we immediately connect them and so  $j$  gets nothing (hence will not be violated) while  $i$  just pays for the distance to go to  $j$ .

In all, throughout Algorithm 7,  $\alpha$  is feasible. ■

**Note (Trim down).** Just like Algorithm 6, after getting the initial solution  $O$ , we'll soon see in the analysis section that it's kind of wasteful, so we trim it down to obtain a better solution.

### 3.2.2 Analysis

We first do a naive analysis, i.e., try to bound the connected cost and opening cost for  $O$  obtained in Algorithm 7 before line 12, which turns out to be not working. The problem is not on connected cost, since as noted above,  $d(i, j) \leq \alpha_i$  so the connection cost is at most  $\sum_i \alpha_i$ .

**Remark.** Bound the opening cost naively can't guarantee a constant approximation factor.

**Proof.** To bound opening cost, we see that

$$\text{opening cost} = f |O| = \sum_{j \in O} f = \sum_j \sum_i \beta_{ij} = \sum_i \sum_{j \in O} \beta_{ij}.$$

Observe that since  $\beta_{ij} = \max(0, \alpha_i - d(i, j)) \leq \alpha_i$ , hence if we can guarantee for each  $i$ , it only pays for one  $j$ , then we will get a 2-approximation. But this might not be the case since we don't have control of how many  $j$  that  $i$  is paying.  $\circledast$

Let's first introduce some notions in order to analyze Algorithm 7.

**Notation** (Connecting witness). The first open facility connected to  $i$  is called the *connecting witness*  $w(i) \in Q$  for every  $i \in P$ .

**Notation** (Contributing). We say  $(i, j)$  is *contributing* if  $\alpha_i > d(i, j)$ , i.e.,  $\beta_{ij} > 0$ .<sup>a</sup>

<sup>a</sup>We now have a strict inequality, i.e.,  $i$  is now paying some non-trivial amount to  $j$ .

Note that the problem in the naive solution happens when a client  $i$  pays multiple facilities  $j$ . And a simple idea is to close some facilities  $j$  such that every client pay at most 1 facility.

**Intuition.** If  $i$  is *contributes* to two facilities  $j$  and  $j'$ , we close down one of them basically since this is where the problem comes from. This is exactly how we *trim down*  $O$ : by considering  $G = (O, E)$  such that  $(j, j') \in E$  if and only if  $\exists i \in P$  that *contributes* to both  $j$  and  $j'$ , taking *maximal independent set* of  $G$  exactly makes  $i$  paying to only one  $j$ .

**Note.** In this case, we take care of opening cost, but the connected cost might be worse, so we basically turn to bound another quantity while still keep one term simple to bound.

**Notation** (Directed connected). We say  $i \in P$  is *directed connected* if  $j \in O'$  such that  $(i, j)$  is connected ( $\alpha_i \geq d(i, j)$ ). For these  $i$ , divide  $\alpha_i$  into  $\alpha_i^f := \beta_{ij}$  and  $\alpha_i^c := d(i, j)$ , i.e.,  $\alpha_i = \alpha_i^f + \alpha_i^c$ .

**Notation** (Indirected connected). We say  $i$  is *indirectly connected* if  $i$  is not *directed connected*,<sup>a</sup> and like in *directed connected*,  $\alpha_i =: \alpha_i^f + \alpha_i^c$  where  $\alpha_i^f = 0$ ,  $\alpha_i^c = \alpha_i$ .

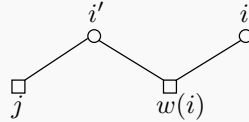


Figure 3.1: When  $i$  is indirect connected.

<sup>a</sup>i.e., there exists  $j$  such that  $(j, w(i)) \in E$ , hence there exists  $i'$  such that  $(i', j)$  and  $(i', w(i))$  *contributing*.

Now, we can bound the opening cost  $f |O'|$  for  $O'$  more carefully. It's now

$$f |O'| = \sum_{j \in O'} \sum_i \beta_{ij} = \sum_{j \in O'} \sum_{\text{d.c. } i} \beta_{ij} = \sum_{\text{d.c. } i} \left[ \sum_{j \in O'} \beta_{ij} \right] = \sum_{\text{d.c. } i} \alpha_i^f.$$

As for connected cost, we see that if  $i$  is **directed connected**,  $d(i, j) \leq \alpha_i^c$ , while if  $i$  is **indirected connected**, it's not so clear. However, we have the following.

**Claim.** If  $i$  is **indirected connected**, then  $d(i, j) \leq 3\alpha_i$ .

**Proof.** Note that  $(j, w(i)) \in E$  and  $d(i, j) \leq \alpha_i + 2\alpha_{i'}$  by looking at [Figure 3.1](#), hence it's sufficient to prove  $\alpha_{i'} \leq \alpha_i$ . To do this, for some facility  $\ell$ , define  $t_\ell$  to be the time  $\ell$  open in [line 6](#), and  $\alpha_i$  be the time  $i$  connected in [line 8](#). We see that

- If  $(i, \ell)$  are **contributing**, then  $\alpha_i \leq t_\ell$ .
- If  $\ell = w(i)$ , then  $t_\ell \leq \alpha_i$ .

Combining these together, we have  $\alpha_{i'} \leq t_{w(i)} \leq \alpha_i$ . ⊛

Finally, we have the following.

**Theorem 3.2.1.** [Algorithm 7](#) is a 3-approximation algorithm.

**Proof.** The cost of  $O'$  produce by [Algorithm 7](#) is just the connected cost of plus the opening cost of  $O'$ , which can be bounded as

$$\text{final cost} = \text{connected cost} + \text{opening cost} \leq \sum_i 3\alpha_i^c + \sum_i \alpha_i^f \leq 3 \sum_i \alpha_i \leq 3 \text{OPT},$$

which shows that it is a 3-approximation algorithm. ■

**Note.** Notice that in the above proof, since we know that the opening cost is exactly  $\sum_i \alpha_i^f$ , and hence even if we pay 3 times of the opening cost, we still get a 3-approximation algorithm.

**Remark.** [Algorithm 7](#) is a very basic algorithm which can be used even as a black-box for other **clustering problems**. We'll revisit this later and consider other **metrics** and see what can we improve.

## Lecture 6: Facility Location with LMP Approximation

### 3.2.3 Hardness

19 Sep. 10:30

For [Problem 3.2.1](#), we have the following.

- 1.488-approximation [[Li13](#)]
- 1.463-approximation is NP-hard [[GK99](#)]

In terms of **LMP**, For [Problem 3.2.1](#), we have the following.

- 3-**LMP** approximation [[JV01](#)]
- 2-**LMP** approximation [[JMS02](#)]
- 1.99...9-**LMP** approximation [[Coh+22](#)]
- 1.73-**LMP** approximation<sup>1</sup> is NP-hard [[JMS02](#)]

### 3.2.4 Greedy Method

Let's take another look at [Problem 3.2.1](#) and see it as an instance of [Problem 2.1.1](#) where the universe is all the clients  $P$ , while the collection of sets are pairs of facility and its connected clients, i.e., **clusters**. Then, it's natural to consider using a similar algorithm as [Algorithm 1](#) to solve this.

<sup>1</sup>The number comes from  $1 + 2/e$ .



**Algorithm 8:** Facility location – Greedy**Data:** A set of clients  $P \subseteq X$ , a set of (possible) facilities  $Q \subseteq X$ , facility cost  $f^a$ **Result:** A set of opened facilities  $O \subseteq Q$ 

```

1  $S \leftarrow \emptyset, O \leftarrow \emptyset$ 
2 while  $S \neq P$  do
3   choose  $(j, T) \in Q \times \mathcal{P}(P \setminus X)$  with minimum  $c((j, T))/|T|$ 
4    $O \leftarrow O \cup \{j\}$ 
5    $S \leftarrow S \cup T$ 
6 return  $O$ 

```

<sup>a</sup>We didn't use it explicitly in the algorithm since we hide it in the cost function  $c(\cdot)$ .

This is just [Algorithm 1](#), hence we have  $H_n$ -approximation. But as we have seen in [Theorem 3.2.1](#), we have achieved a constant approximation ratio for [Problem 3.2.1](#). Hence, we should be able to do better based on [Algorithm 8](#).

**Remark.** If we modify [Algorithm 8](#) such that for all  $(j, T)$ , if  $j$  is open, then we define the cost of this cluster as

$$c((j, T)) := \frac{\sum_{i \in T} d(i, j)}{|T|}.$$

We'll achieve 1.861-approximation, but the analysis is complex.

Instead, we're going to see other variations based on [Algorithm 8](#).

**First Modification**

We see observe that  $c((j, T))/|T|$  is increasing in [Algorithm 8](#). Also, if  $\alpha := c((j, T))/|T|$ , then for all  $i \in T$ ,  $d(i, j) \leq \alpha$  where we interpret this as  $i$  pays  $\alpha_i$  to cover the connection cost  $d(i, j)$  and the opening cost  $\alpha_i - d(i, j)$  of  $j$ . Following this intuition, if we change [line 6](#) in [Algorithm 7](#) (with only first phase) such that the summation is over  $P \setminus S$ , it becomes exactly [Algorithm 8](#).

proof!

**Algorithm 9:** Facility location – Greedy Modification I**Data:** A set of clients  $P \subseteq X$ , a set of (possible) facilities  $Q \subseteq X$ , facility cost  $f$ **Result:** A set of opened facilities  $O \subseteq Q$ 

```

1  $S \leftarrow \emptyset, O \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $O$ :open facilities
2
3 while  $S \neq P$  do
4   while True do
5     increase all  $\{\alpha_i\}_{i \in P \setminus S}$  by a unit
6     if some  $j \in Q \setminus O$  s.t.  $\sum_{i \in P \setminus S} \beta_{ij} = f$  then //  $j$  gets tight (open)
7       break
8     else if some  $i \in P \setminus S$  s.t.  $\alpha_i \geq d(i, j)$  then //  $i$  can connect to  $j \in O$ 
9       break
10     $O \leftarrow O \cup \{j\}$  // Update  $O$ 
11     $S \leftarrow S \cup \{i \in P \setminus S : \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $O$ 

```

**Remark.** Since [line 6](#) and [line 8](#) can happen simultaneously, while what we just said assumes the opposite, so we need to further modify [Algorithm 7](#) in [line 10](#) and [line 11](#)

**Second Modification**

Another potential modification gives us a 1.61-approximation. We essentially allow  $i \in S$  to *switch* in [Algorithm 9](#), i.e., after  $i$  connects to  $j$ , if  $j'$  is closer to  $i$  later,  $i$  can *offer* with  $d(i, j) - d(i, j')$  to other facilities.

**Algorithm 10: Facility location – Greedy Modification II****Data:** A set of clients  $P \subseteq X$ , a set of (possible) facilities  $Q \subseteq X$ , facility cost  $f$ **Result:** A set of opened facilities  $O \subseteq Q$ 

```

1  $S \leftarrow \emptyset, O \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $O$ :open facilities
2
3 while  $S \neq P$  do
4   while True do
5     increase all  $\{\alpha_i\}_{i \in P \setminus S}$  by a unit
6     if some  $j \in Q \setminus O$  s.t.  $\sum_{i \in S} (d(i, w(i)) - d(i, j))^+ + \sum_{i \in P \setminus S} \beta_{ij} = f^a$  then
7       break
8     if some  $i \in P \setminus S$  s.t.  $\alpha_i \geq d(i, j)$  then //  $i$  can connect to  $j \in O$ 
9       break
10     $O \leftarrow O \cup \{j\}$  // Update  $O$ 
11     $S \leftarrow S \cup \{i \in P \setminus S : \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $O$ 

```

<sup>a</sup>We define  $a^+ := \max(a, 0)$  and also,  $w(i)$  is now the *current* facility  $i$  is connected to.**Third Modification**

If we run Algorithm 10 with facility cost being  $\hat{f} := 2f$ , we can have a 2-LMP approximation algorithm as follows.

**Algorithm 11: Facility location – Greedy Modification III****Data:** A set of clients  $P \subseteq X$ , a set of (possible) facilities  $Q \subseteq X$ , facility cost  $f$ **Result:** A set of opened facilities  $O \subseteq Q$ 

```

1  $S \leftarrow \emptyset, O \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $O$ :open facilities
2
3 while  $S \neq P$  do
4   while True do
5     increase all  $\{\alpha_i\}_{i \in P \setminus S}$  by a unit
6     if some  $j \in Q \setminus O$  s.t.  $\sum_{i \in S} (d(i, w(i)) - d(i, j))^+ + \sum_{i \in P \setminus S} \beta_{ij} = \hat{f}^a$  then
7       break
8     if some  $i \in P \setminus S$  s.t.  $\alpha_i \geq d(i, j)$  then //  $i$  can connect to  $j \in O$ 
9       break
10     $O \leftarrow O \cup \{j\}$  // Update  $O$ 
11     $S \leftarrow S \cup \{i \in P \setminus S : \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $O$ 

```

<sup>a</sup>We define  $a^+ := \max(a, 0)$  and also,  $w(i)$  is now the *current* facility  $i$  is connected to.

It's clear that in Algorithm 10, the connection cost plus 2 times the opening cost is  $\sum_{i \in P} \alpha_i$  from how we design the algorithm by changing the facility cost from  $f$  to  $\hat{f} := 2f$ . Now, a crucial lemma is the following.

**Lemma 3.2.2.**  $(\alpha', \beta')$  is dual feasible, where  $\alpha' := \alpha/2$ ,  $\beta'_{ij} := (\alpha'_{ij} - d(i, j))^+$ .

**Proof.** It's sufficient to consider  $j \in Q$  and prove that  $\sum_{i \in P'} \beta'_{ij} \leq f$  where  $P' := \{i : \beta'_{ij} > 0\} = [n]$  where we're overloading  $n$  here. Let's order  $\alpha_i$  such that  $\alpha_1 \leq \dots \leq \alpha_n$  where  $\alpha_i$  is the time  $i$  when  $i$  is *first connected*.

**Claim.** For all  $i, k \in P'$  such that  $\alpha_k \leq \alpha_i$ , at time (right before)  $\alpha_i$ , offer from  $k$  to  $j^a$  is at most  $\alpha_i - d(i, j) - 2d(k, j)$  for any  $j \in Q$ .

<sup>a</sup>We assume  $k$  currently (or is going to) connects to  $j'$ .

**Proof.** We see that if  $\alpha_i = \alpha_k$ , the offer is just  $(\alpha_k - d(i, j))^+$ . Otherwise, we have  $\alpha_k < \alpha_i$ . If  $\alpha_i > d(k, j') + d(k, j) + d(i, j)$ , we immediately get a contradiction since from triangle inequality,  $\alpha_i > d(i, j')$ , i.e.,  $i$  already connect to  $j'$ . Hence,

$$\alpha_i \leq d(k, j') + d(k, j) + d(i, j).$$

Then, the offer from  $k$  to  $j$  is  $(d(k, j') - d(k, j))^+ \geq \alpha_i - d(i, j) - 2d(k, j)$ . ⊗

Observe that for all  $i \in [n]$ , we have

$$\sum_{k=1}^{i-1} (\alpha_i - d(i, j) - 2d(k, j)) + \sum_{k=i}^n (\alpha_i - d(k, j)) \leq \hat{f} \quad (3.1)$$

by considering the total offer from  $k$  to  $j$  at time (right before)  $\alpha_i$ . Now, we add Equation 3.1 for all  $i \in [n]$ , we have

$$n \sum_{i=1}^n \alpha_i - \sum_{i=1}^n (i-1)d(i, j) - 2 \sum_{k=1}^n (n-k)d(k, j) - \sum_{i=1}^n k \cdot d(k, j) \leq n\hat{f} = 2nf.$$

Since the summation over  $k$  is just indexes, we can change it to  $i$ , hence

$$\begin{aligned} 2nf &\geq n \sum_{i=1}^n \alpha_i - \sum_{i=1}^n (i-1)d(i, j) - 2 \sum_{i=1}^n (n-i)d(i, j) - \sum_{i=1}^n i \cdot d(i, j) \\ &\geq n \sum_{i=1}^n \alpha_i - \sum_{i=1}^n id(i, j) - 2 \sum_{i=1}^n (n-i)d(i, j) - \sum_{i=1}^n i \cdot d(i, j) = n \sum_{i=1}^n \alpha_i - \sum_{i=1}^n 2nd(i, j) \end{aligned}$$

where we turn the factor  $(i-1)$  into  $i$  and gather the terms together. Clean up a bit, we have

$$n \sum_{i=1}^n \alpha_i - 2n \sum_{i=1}^n d(i, j) \leq 2nf \Leftrightarrow \frac{\sum_{i=1}^n \alpha_i}{2} - \sum_{i=1}^n d(i, j) \leq f,$$

finishing the proof. ■

From Lemma 3.2.2, we immediately have the following.

**Theorem 3.2.2.** Algorithm 10 is a 2-LMP approximation algorithm w.r.t. the original  $f$ .

## Lecture 7: $k$ -Median and LMP Approximation

### 3.3 $k$ -Median

21 Sep. 10:30

Let's look at another clustering problem.

**Problem 3.3.1 ( $k$ -median).** Given a metric space  $(X, d)$  and  $P, Q \subseteq X$  with  $k \in \mathbb{N}$ , find  $Q' \subseteq Q$  with  $|Q'| = k$  which minimizes  $\sum_{i \in P} \min_{j \in Q'} d(i, j)$ .

The natural linear programming is the following. Consider  $\{x_{ij}\}_{i \in P, j \in Q}$  and  $\{y_j\}_{j \in Q}$ , then

$$\begin{aligned}
 \min \quad & \sum_{ij} x_{ij} d(i, j) \\
 \sum_j x_{ij} & \geq 1 & \forall i \in P & (\alpha_i) \\
 x_{ij} & \leq y_j & \forall i \in P, j \in Q & (\beta_{ij}) \\
 \sum_j y_j & \leq k \\
 x, y & \geq 0
 \end{aligned}$$

The dual is then

$$\begin{aligned}
 \max \quad & \sum_i \alpha_i - kf \\
 \sum_i \beta_{ij} & \leq d(i, j) & \forall i \in P, j \in Q \\
 \sum_i \beta_{ij} & \leq f & \forall j \in Q \\
 \alpha, \beta & \geq 0
 \end{aligned}$$

**Definition 3.3.1 (LMP approximation).** An algorithm is called  $\gamma$ -Lagrangian multiplier preserving approximation (LMP-approximation) if

$$\frac{\text{conn}(\text{ALG})}{\gamma} + k'f \leq \sum_i \alpha_i$$

for some  $\gamma > 0$ .

Fix

**Remark.** Suppose we guessed  $f$ , and after running  $\gamma$ -LMP approximation algorithm and somehow get  $k' = k$ . Then we have

$$\frac{\text{conn}(\text{ALG})}{\gamma} \leq \sum_i \alpha_i - kf \leq \text{OPT}_{k\text{-med.}},$$

i.e., this is a  $\gamma$ -approximation algorithm.

Actually, we don't have ideas about the relation between  $k$  and  $f$ . In particular, we have very little knowledge about which, the only thing we know is it behaves quite arbitrary. In this case, our previous technique doesn't work anymore, a new idea is then to maintain  $[f^2, f^1]$  such that  $f^2 \leq f^1$ , where

- At  $f^2$ , the algorithm opens  $k^2 \geq k$  facilities.
- At  $f^1$ , the algorithm opens  $k^1 \leq k$  facilities.

Then, by binary search, we can get  $f_2 \leq f_1$  such that

$$|f^1 - f^2| \leq \frac{\epsilon \text{OPT}}{n}.$$

### 3.3.1 Bipoint Rounding

Now, if  $a \in [0, 1]$  and  $b := 1 - a$ , we have  $k := ak^1 + bk^2$  where  $k^1 \leq k \leq k^2$ . Denote  $C^i$  as the connection cost with  $f^i$  such that  $C^1 \geq C^2$ , we have

$$\begin{cases} C^1 + \gamma k^1 f^1 \leq \gamma \sum_i \alpha_i^1, & (\times a) \\ C^2 + \gamma k^2 f^2 \leq \gamma \sum_i \alpha_i^2, & (\times b) \end{cases}$$

hence,

$$aC^1 + bC^2 \leq \gamma \left( a \sum_i \alpha_i^1 + b \sum_i \alpha_i^2 - ak^1 f^1 - bk^2 f^2 \right) \leq \underbrace{\gamma \left( \sum_i \alpha_i - kf \right)}_{\leq \text{OPT}_{k\text{-med.}}} + \underbrace{\gamma k |f^1 - f^2|}_{\in \text{OPT}_{k\text{-med.}}},$$

where we set  $\alpha := a\alpha^1 + b\alpha^2$  and  $f := \max(f^1, f^2)$ .

**Note.**  $(\alpha, f)$  is dual-feasible for [Problem 3.3.1](#).

**Definition 3.3.2 (Bipoint solution).** Given  $F^1, F^2$  with  $|F^1| = k^1, |F^2| = k^2$  and  $k = ak^1 + bk^2$  for  $a, b \in [0, 1], a + b = 1$ . If the connection cost of  $aF^1 + bF^2$  satisfies

$$a \cdot \text{conn}(F^1) + b \cdot \text{conn}(F^2) \leq \gamma \cdot \text{OPT}_{k\text{-med.}},$$

we then call such a solution a *bipoint solution*.

**Remark ( $\delta$ -bipoint rounding).** For  $F^1$  and  $F^2$ , output  $F$  with  $|F| = k$  such that

$$\text{conn}(F) \leq \delta \cdot (aC^1 + bC^2).$$

## 2-Bipoint Rounding

For  $j \in F^1$ , let  $\pi(j)$  be the closest facility in  $F^2$  to  $j$ . Let  $F^* = \{j' \in F^2 : j' = \pi(j) \text{ for some } j \in F^1\}$ . We see that  $|F^*| \leq k^1$ . If  $|F^*| < k^1$ , then we add arbitrary centers so that  $|F^*| = k^1$ .

To open the facilities as what we want, consider the following rounding algorithm.

1. Open  $F^1$  with probability  $a$ , otherwise open  $F^*$ .
2. Randomly choose  $k - k_1$  far from  $F^2 \setminus F^*$ .

Then,

1.  $j \in F^1, \Pr(j \text{ open}) = a$
2.  $j \in F^*, \Pr(j \text{ open}) = b$
3.  $j \in F^2 \setminus F^*, \Pr(j \text{ open}) = \frac{k - k_1}{k_2 - k_1} = b$

Now, fix  $i \in P$ , we have

	distance	probability
$j^2$ open	$d_2$	$b$
$j^2$ not open, $j^1$ open	$d_1$	$\geq (a - b)^+ =: M$
none of $j^1, j^2$ open	$2d_1 + d_2$	$\leq 1 - b - M$

Then, the expected cost is just

$$\mathbb{E}[i\text{'s connected cost}] \leq bd_2 + Md_1 + (1 - b - M)(2d_1 + d_2).$$

- If  $b \geq a$ ,<sup>2</sup> then  $M = 0$  and

$$\mathbb{E}[i\text{'s connected cost}] \leq b - d_2 + (1 - b)(2d_1 + d_2) = 2ad_1 + d_2 \leq 2(ad_1 + bd_2).$$

- If  $a > b$ , then  $a > 1/2$ ,  $M = a - b$  and

$$\mathbb{E}[i\text{'s connected cost}] \leq bd_2((a - b)d_1 + b(2d_1 + d_2)) + d_1(a + b) + d_2(b + d) \leq 2(ad_1 + bd_2).$$

<sup>2</sup>We have  $b \geq 1/2$ .

---

**Remark (SOTA).** The SOTA result

# Appendix

# Bibliography

- [BS14] Boaz Barak and David Steurer. *Sum-of-squares proofs and the quest toward optimal algorithms*. 2014. DOI: [10.48550/ARXIV.1404.5236](https://arxiv.org/abs/1404.5236). URL: <https://arxiv.org/abs/1404.5236>.
- [Coh+22] Vincent Cohen-Addad et al. *Breaching the 2 LMP Approximation Barrier for Facility Location with Applications to k-Median*. 2022. DOI: [10.48550/ARXIV.2207.05150](https://arxiv.org/abs/2207.05150). URL: <https://arxiv.org/abs/2207.05150>.
- [GK99] Sudipto Guha and Samir Khuller. “Greedy Strikes Back: Improved Facility Location Algorithms”. In: *Journal of Algorithms* 31.1 (1999), pp. 228–248. ISSN: 0196-6774. DOI: <https://doi.org/10.1006/jagm.1998.0993>. URL: <https://www.sciencedirect.com/science/article/pii/S0196677498909932>.
- [JMS02] Kamal Jain, Mohammad Mahdian, and Amin Saberi. “A New Greedy Approach for Facility Location Problems”. In: STOC ’02. Montreal, Quebec, Canada: Association for Computing Machinery, 2002, pp. 731–740. ISBN: 1581134959. DOI: [10.1145/509907.510012](https://doi.org/10.1145/509907.510012). URL: <https://doi.org/10.1145/509907.510012>.
- [JV01] Kamal Jain and Vijay V. Vazirani. “Approximation Algorithms for Metric Facility Location and K-Median Problems Using the Primal-Dual Schema and Lagrangian Relaxation”. In: *J. ACM* 48.2 (Mar. 2001), pp. 274–296. ISSN: 0004-5411. DOI: [10.1145/375827.375845](https://doi.org/10.1145/375827.375845). URL: <https://doi.org/10.1145/375827.375845>.
- [Li13] Shi Li. “A 1.488 approximation algorithm for the uncapacitated facility location problem”. In: *Information and Computation* 222 (2013). 38th International Colloquium on Automata, Languages and Programming (ICALP 2011), pp. 45–58. ISSN: 0890-5401. DOI: <https://doi.org/10.1016/j.ic.2012.01.007>. URL: <https://www.sciencedirect.com/science/article/pii/S0890540112001459>.
- [ODo21] Ryan O’Donnell. *Analysis of Boolean Functions*. 2021. DOI: [10.48550/ARXIV.2105.10386](https://arxiv.org/abs/2105.10386). URL: <https://arxiv.org/abs/2105.10386>.
- [Vaz02] V.V. Vazirani. *Approximation Algorithms*. Springer Berlin Heidelberg, 2002. ISBN: 9783540653677. URL: <https://books.google.com/books?id=EILqAmzKgYIC>.
- [WS11] D.P. Williamson and D.B. Shmoys. *The Design of Approximation Algorithms*. Cambridge University Press, 2011. ISBN: 9781139498173. URL: [https://books.google.com/books?id=Cc%5C\\_Fdqf3bBgC](https://books.google.com/books?id=Cc%5C_Fdqf3bBgC).