

EECS598-001

Approximation Algorithms & Hardness of Approximation

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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

1	Introduction	2
1.1	Computational Problem	2
1.2	Efficient Algorithms	2
1.3	Approximation Algorithms	3
1.4	Hardness	4
2	Covering	5
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	k -Median	23
2.7	Euclidean k -Median	30
3	Traveling Salesman Problem	34
3.1	Spanning Tree	34
3.2	Negative Correlation	38
3.3	Asymmetric Traveling Salesman Problem	39
3.4	Symmetric Traveling Salesman Problem	46
3.5	Beyond the $3/2$ Barrier for STSP	48
4	Semidefinite Programming and Lasserre Hierarchy	54
4.1	Semidefinite Programming	54
4.2	Lasserre Hierarchy	57
4.3	Graph Coloring	63
5	Hardness of Approximation	68
5.1	Approximation Complexity	68
5.2	Probabilistically Checkable Proofs	70
5.3	FGLSS Graph	72
5.4	Label Cover	74
A	Review	77
A.1	Boolean Satisfaction Problem	77

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 α -*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 α 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 α -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 Ω (often called *universe*), the *set system* is an order pair (Ω, \mathcal{S}) where \mathcal{S} is a collection of subsets of Ω .

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

Definition 2.1.2 (Degree). Given a [set system](#) (Ω, \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 (Ω, \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 2.1: 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 2.1](#) is a polynomial time algorithm, also, the output \mathcal{S}' is always a valid [covering](#).

Theorem 2.2.1. [Algorithm 2.1](#) is an H_k -approximation^a algorithm.

^a 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 2.1](#) 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 2.1](#) 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 LP¹ 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

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

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

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

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 2.2: **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 2.2**'s approximation ratio.

¹See **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 2.2 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 OPT is assumed to be OPT_{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 OPT and OPT_{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 2.2 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 2.3: 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 2.4: Set cover – Multi-time Randomized LP Rounding

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

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 2.4 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 2.4 is correct with probability at least $1 - 1/n$.

Lemma 2.3.4. With $\alpha = 2 \ln n$, \mathcal{S}' returned from Algorithm 2.4 has an approximation ratio $4 \ln n$ w.p. at least $\frac{1}{2}$.^a

^aNote 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 2.4 many times, we get a $(4 \ln n)$ -approximation algorithm with high probability.^a

^aNote 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 2.4 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 2.4](#), 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 $\text{OPT}_{\text{SC}} \geq |C|$.^a

^aSC 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 $\text{OPT}_{\text{SIS}} \leq \text{OPT}_{\text{SC}}$.^a

^aSIS 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.^a ■

^aNote 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 2.1 while maintaining y_e for every e , and we're done.

Theorem 2.4.1. Algorithm 2.1 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 OPT_{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/α is dual-feasible.
3. This shows the algorithm is α -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 2.5: [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 2.5](#) is correct and can be implemented efficiently.

Theorem 2.5.1. [Algorithm 2.5](#) 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 6: Facility Location with LMP Approximation

2.5.1 Hardness

19 Sep. 10:30

For ??, we have the following.

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

Turns out that specifically for ??, we can have a more refine notion of approximation ratio defined below.

Definition 2.5.1 (LMP approximation). An algorithm ALG which solves facility location is called γ -Lagrangian multiplier preserving approximation (LMP-approximation) if

$$\frac{\text{conn}(\text{ALG})}{\gamma} + \text{open}(\text{ALG}) \leq \sum_i \alpha_i$$

for some $\gamma > 0$.^a

^aThe opening cost is just $k'f$ if ALG opens k' centers.

Remark. The notion of [LMP approximation](#) is due to Lagrangian multiplier in the field of optimization, where the dual variables are treated as a Lagrangian multipliers. And [Definition 2.5.3](#) says that we're not approximating $k'f$ at all, hence it's *preserving*.

And indeed, we now have a more refined characterization about ??.

Corollary 2.5.1. ?? is a 3-[LMP approximation](#) algorithm.

Remark (SOTA). If we look at the SOTA result in terms of LMP, we have the following.

- (a) 3-LMP approximation [JV01]
- (b) 2-LMP approximation [JMS02]
- (c) $1.99 \dots 9$ -LMP approximation [Coh+22]
- (d) 1.73-LMP approximation^a is NP-hard [JMS02]

^aThe number comes from $1 + 2/e$.

2.5.2 Greedy Method

Let's take another look at ?? 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 2.1 to solve this.

Algorithm 2.6: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset$ 
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    $Q' \leftarrow Q' \cup \{j\}$ 
5    $S \leftarrow S \cup T$ 
6 return  $Q'$ 
```

^aWe didn't use it explicitly in the algorithm since we hide it in the cost function $c(\cdot)$.

This is just Algorithm 2.1, hence we have H_n -approximation. But as we have seen in ??, we have achieved a constant approximation ratio for ?. Hence, we should be able to do better based on Algorithm 2.14.

Remark. If we modify Algorithm 2.14 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 2.14.

First Modification

We see observe that $c((j, T))/|T|$ is increasing in Algorithm 2.14. 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 α_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 ?? in ?? (with only first phase) such that the summation is over $P \setminus S$, it becomes exactly Algorithm 2.14.

proof!

Algorithm 2.7: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11     $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

Remark. Since ?? and ?? can happen simultaneously, while what we just said assumes the opposite, so we need to further modify ?? in ?? and ??.

Second Modification

Another potential modification gives us a 1.61-approximation. We essentially allow $i \in S$ to *switch* in [Algorithm 2.15](#), 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 2.8: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11     $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

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

Third Modification

If we run [Algorithm 2.16](#) with facility cost being $\hat{f} := 2f$, we can have a 2-LMP approximation algorithm as follows.

Algorithm 2.9: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11     $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

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

It's clear that in [Algorithm 2.17](#), 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 2.5.1. (α', β') is dual feasible, where $\alpha' := \alpha/2$, $\beta'_{ij} := (\alpha'_i - d(i, j))^+$.

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 α_i such that $\alpha_1 \leq \dots \leq \alpha_n$ where α_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) α_i , offer from k to j^a is at most $\alpha_i - d(i, j) - 2d(k, j)$ for any $j \in Q$.

^aWe 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

$$\sum_{k=1}^{i-1} (\alpha_i - d(i, j) - 2d(k, j)) + \sum_{k=i}^n (\alpha_i - d(k, j)) \leq \hat{f} \quad (2.1)$$

by considering the total offer from k to j at time (right before) α_i . Now, we add [Equation 2.3](#) 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 2.5.3](#), we immediately have the following.

Theorem 2.5.2. [Algorithm 2.17](#) is a 2-LMP approximation algorithm w.r.t. the original f .

Lecture 6: Facility Location with LMP Approximation

2.5.3 Hardness

19 Sep. 10:30

For ??, we have the following.

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

Turns out that specifically for ??, we can have a more refine notion of approximation ratio defined below.

Definition 2.5.2 (LMP approximation). An algorithm ALG which solves facility location is called γ -Lagrangian multiplier preserving approximation (LMP-approximation) if

$$\frac{\text{conn}(\text{ALG})}{\gamma} + \text{open}(\text{ALG}) \leq \sum_i \alpha_i$$

for some $\gamma > 0$.^a

^aThe opening cost is just $k'f$ if ALG opens k' centers.

Remark. The notion of [LMP approximation](#) is due to Lagrangian multiplier in the field of optimization, where the dual variables are treated as a Lagrangian multipliers. And [Definition 2.5.3](#) says that we're not approximating $k'f$ at all, hence it's *preserving*.

And indeed, we now have a more refined characterization about ??.

Corollary 2.5.2. ?? is a 3-LMP approximation algorithm.

Remark (SOTA). If we look at the SOTA result in terms of [LMP](#), we have the following.

- (a) 3-LMP approximation [[JV01](#)]
- (b) 2-LMP approximation [[JMS02](#)]
- (c) $1.99\dots 9$ -LMP approximation [[Coh+22](#)]
- (d) 1.73-LMP approximation^a is NP-hard [[JMS02](#)]

^aThe number comes from $1 + 2/e$.

2.5.4 Greedy Method

Let's take another look at ?? 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 2.1](#) to solve this.

Algorithm 2.10: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset$ 
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    $Q' \leftarrow Q' \cup \{j\}$ 
5    $S \leftarrow S \cup T$ 
6 return  $Q'$ 

```

^aWe didn't use it explicitly in the algorithm since we hide it in the cost function $c(\cdot)$.

This is just [Algorithm 2.1](#), hence we have H_n -approximation. But as we have seen in ??, we have achieved a constant approximation ratio for ??. Hence, we should be able to do better based on [Algorithm 2.14](#).

Remark. If we modify [Algorithm 2.14](#) 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 2.14](#).

First Modification

We see observe that $c((j, T))/|T|$ is increasing in [Algorithm 2.14](#). 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 α_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 ?? in ?? (with only first phase) such that the summation is over $P \setminus S$, it becomes exactly [Algorithm 2.14](#).

proof!

Algorithm 2.11: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11     $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

Remark. Since ?? and ?? can happen simultaneously, while what we just said assumes the opposite, so we need to further modify ?? in ?? and ??.

Second Modification

Another potential modification gives us a 1.61-approximation. We essentially allow $i \in S$ to *switch* in [Algorithm 2.15](#), 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 2.12: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11     $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

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

Third Modification

If we run [Algorithm 2.16](#) with facility cost being $\hat{f} := 2f$, we can have a 2-LMP approximation algorithm as follows.

Algorithm 2.13: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11     $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

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

It's clear that in [Algorithm 2.17](#), 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 2.5.2. (α', β') is dual feasible, where $\alpha' := \alpha/2$, $\beta'_{ij} := (\alpha'_i - d(i, j))^+$.

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 α_i such that $\alpha_1 \leq \dots \leq \alpha_n$ where α_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) α_i , offer from k to j^a is at most $\alpha_i - d(i, j) - 2d(k, j)$ for any $j \in Q$.

^aWe 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

$$\sum_{k=1}^{i-1} (\alpha_i - d(i, j) - 2d(k, j)) + \sum_{k=i}^n (\alpha_i - d(k, j)) \leq \hat{f} \quad (2.2)$$

by considering the total offer from k to j at time (right before) α_i . Now, we add Equation 2.3 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 2.5.3, we immediately have the following.

Theorem 2.5.3. Algorithm 2.17 is a 2-LMP approximation algorithm w.r.t. the original f .

Lecture 6: Facility Location with LMP Approximation

2.5.5 Hardness

19 Sep. 10:30

For ??, we have the following.

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

Turns out that specifically for ??, we can have a more refine notion of approximation ratio defined below.

Definition 2.5.3 (LMP approximation). An algorithm ALG which solves facility location is called γ -Lagrangian multiplier preserving approximation (LMP-approximation) if

$$\frac{\text{conn}(\text{ALG})}{\gamma} + \text{open}(\text{ALG}) \leq \sum_i \alpha_i$$

for some $\gamma > 0$.^a

^aThe opening cost is just $k'f$ if ALG opens k' centers.

Remark. The notion of **LMP approximation** is due to Lagrangian multiplier in the field of optimization, where the dual variables are treated as a Lagrangian multipliers. And **Definition 2.5.3** says that we're not approximating $k'f$ at all, hence it's *preserving*.

And indeed, we now have a more refined characterization about ??.

Corollary 2.5.3. ?? is a 3-LMP approximation algorithm.

Remark (SOTA). If we look at the SOTA result in terms of **LMP**, we have the following.

- (a) 3-LMP approximation [JV01]
- (b) 2-LMP approximation [JMS02]
- (c) 1.99...9-LMP approximation [Coh+22]
- (d) 1.73-LMP approximation^a is NP-hard [JMS02]

^aThe number comes from $1 + 2/e$.

2.5.6 Greedy Method

Let's take another look at ?? 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 2.1** to solve this.

Algorithm 2.14: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset$ 
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    $Q' \leftarrow Q' \cup \{j\}$ 
5    $S \leftarrow S \cup T$ 
6 return  $Q'$ 
```

^aWe didn't use it explicitly in the algorithm since we hide it in the cost function $c(\cdot)$.

This is just **Algorithm 2.1**, hence we have H_n -approximation. But as we have seen in ??, we have achieved a constant approximation ratio for ??. Hence, we should be able to do better based on **Algorithm 2.14**.

Remark. If we modify **Algorithm 2.14** 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 2.14**.

First Modification

We see observe that $c((j, T))/|T|$ is increasing in [Algorithm 2.14](#). 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 α_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 ?? in ?? (with only first phase) such that the summation is over $P \setminus S$, it becomes exactly [Algorithm 2.14](#).

proof!

Algorithm 2.15: 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 $Q' \subseteq Q$

```

1   $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9              break
10      $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11      $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

Remark. Since ?? and ?? can happen simultaneously, while what we just said assumes the opposite, so we need to further modify ?? in ?? and ??.

Second Modification

Another potential modification gives us a 1.61-approximation. We essentially allow $i \in S$ to *switch* in [Algorithm 2.15](#), 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 2.16: 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 $Q' \subseteq Q$

```

1   $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9              break
10      $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11      $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

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

Third Modification

If we run [Algorithm 2.16](#) with facility cost being $\hat{f} := 2f$, we can have a 2-LMP approximation algorithm as follows.

Algorithm 2.17: 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :connected clients,  $Q'$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow Q' \cup \{j\}$  // Update  $Q'$ 
11     $S \leftarrow S \cup \{i \in P \setminus S: \alpha_i \geq d(i, j)\}$  // Update  $S$ 
12 return  $Q'$ 

```

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

It's clear that in Algorithm 2.17, 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 2.5.3. (α', β') is dual feasible, where $\alpha' := \alpha/2$, $\beta'_{ij} := (\alpha'_i - d(i, j))^+$.

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 α_i such that $\alpha_1 \leq \dots \leq \alpha_n$ where α_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) α_i , offer from k to j^a is at most $\alpha_i - d(i, j) - 2d(k, j)$ for any $j \in Q$.

^aWe 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

$$\sum_{k=1}^{i-1} (\alpha_i - d(i, j) - 2d(k, j)) + \sum_{k=i}^n (\alpha_i - d(k, j)) \leq \hat{f} \quad (2.3)$$

by considering the total offer from k to j at time (right before) α_i . Now, we add Equation 2.3 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 2.5.3](#), we immediately have the following.

Theorem 2.5.4. [Algorithm 2.17](#) is a 2-LMP approximation algorithm w.r.t. the original f .

Lecture 7: k -Median and LMP Approximation

2.6 k -Median

21 Sep. 10:30

Let's look at another clustering problem.

Problem 2.6.1 (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 for [Problem 2.6.1](#) is the following. Consider $\{x_{ij}\}_{i \in P, j \in Q}$ and $\{y_j\}_{j \in Q}$, then

$$\begin{aligned} \min \quad & \sum_{ij} x_{ij} d(i, j) \\ & \sum_j x_{ij} \geq 1 \quad \forall i \in P \quad (\alpha_i) \\ & x_{ij} \leq y_j \quad \forall i \in P, j \in Q \quad (\beta_{ij}) \\ & \sum_j y_j \leq k \quad (f)^2 \\ & x, y \geq 0 \end{aligned}$$

Intuition. We interpret x_{ij} as follows: if $x_{ij} = 1$, then i belongs to j . And $y_j = 1$ if j is the actual median we choose (i.e., in Q'). As for constraints, both $\sum_j x_{ij} \geq 1$ and $\sum_j y_j \leq k$ are clear, while for $x_{ij} \leq y_j$, we see that it can't be the case that $x_{ij} = 1$ while $y_j = 0$, i.e., we can't have the case that x_{ij} belongs to j while j isn't even in Q' .

The dual is then

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

²Notice that compare to ??, f here is a variable but not a given facility cost! The reason why we do this will be clear soon.

Note. Notice that this is exactly the dual as ??, except that we now have an additional $-kf$ term in the objective function. Although f is not included in the statement of Problem 2.6.1, by denoting one of the dual variable f , we get a similar formulation compare to ??.

Due to the similarity between Problem 2.6.1 and ??, we can try to use Algorithm 2.17 which solves ?? with 2-LMP guarantee. But note that in ??, we need to specify f . Suppose we guessed f , and we run a γ -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 γ -approximation algorithm. So now, the task is to guess f such that the algorithm gives exactly k centers.

2.6.1 Bipoint Solution

Turns out that we don't have ideas about the relation between k and f , the only thing we know is that if $f \rightarrow \infty$, k decreases, other than that it behaves quite arbitrary.

Remark. The relation between k and f indeed highly depends on what algorithm we use. But at least for Algorithm 2.17, nobody knows anything in this case.

Given this fact, just randomly guess one f doesn't work. A new idea is then to maintain two solutions (or interval) $[f^2, f^1]$ such that $f^2 \leq f^1$,³ where

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

Then, a naive approach is to use binary search and get $f^2 \leq f^1$ such that

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

Notice that the whole point of doing binary search is because we assume that if $k^2 \geq k$ at f^2 and $k^1 \leq k$ at f^1 , then we can find an $f^* \in [f^2, f^1]$ such that we get exactly $k^* = k$ at f^* .

Remark (Caveat of achieving k). This is probably not the case for Algorithm 2.14 (2-LMP) since the decision is quite sequential; but if we use ?? (2-LMP), since there are lots of maximal independent sets, so by doing **a lot more work**, we can actually achieve this.

Now, assume that we have continuity of the relation between k and f by carefully designing our (γ -LMP) algorithm, then $\exists a \in [0, 1]$ and $b := 1 - a$ such that $k := ak^1 + bk^2$ where $k^1 \leq k \leq k^2$. Denote C^i as the connection cost $\text{conn}(f^i)$ of 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,

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

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

³We start from $f^2 = 0$ and $f^1 = \infty$, where we set f^1 arbitrary large.

Note. To make sure $\sum_i \alpha_i - kf \leq \text{OPT}_{k\text{-med.}}$, we need to check that (α, f) is dual-feasible for Problem 2.6.1.

Proof. The feasibility comes from the fact that the first two constraints of Problem 2.6.1 are linear, so they're automatically satisfied. The only non-trivial constraint is $\sum_i \beta_{ij} \leq f$, but since we choose f to be the maximum, it'll be more satisfied. \otimes

Definition 2.6.1 (Bipoint solution). Given F^1, F^2 with $|F^1| = k^1, |F^2| = k^2$ and $k = ak^1 + bk^2$ for $a, b \in [0, 1]$ and $a + b = 1$, the *bipoint solution*, denoted as $aF^1 + bF^2$, satisfies

$$aC^1 + bC^2 \leq \gamma \cdot \text{OPT}_{k\text{-med.}}$$

2.6.2 Bipoint Rounding

From Definition 2.6.1, it's natural to do the so-called *bipoint rounding*.

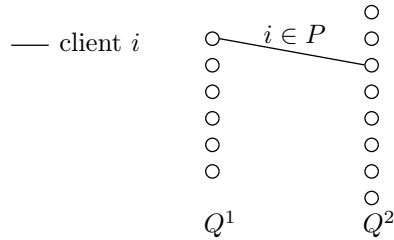
Definition 2.6.2 (δ -bipoint rounding). Given solutions F^1 and F^2 , a solution F with $|F| = k$ such that

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

is the so-called *δ -bipoint rounding solution*.

Note. If we have a δ -bipoint rounding of a γ -LMP algorithm solution, then we automatically have an approximation ratio of $\delta \cdot \gamma$ for this bipoint rounding solution.

Back to Problem 2.6.1, we see that we can actually get a 2-bipoint rounding as follows. Consider we create a bipartite graph with $Q^1, Q^2 \subseteq Q$ being two sides of the graph. Then for each $i \in P$, i is connected to the closest facility in Q^1 , and also another closest facility in Q^2 , so we can create an edge between these two facilities.



Now, for a fixed $i \in P$, let $d_j := d(i, Q^j)$ for $j = 1, 2$, we want to compare our designed final cost to $aC^1 + bC^2$, so for this fixed i , we want to make sure i pays not much more than $ad_1 + bd_2$.

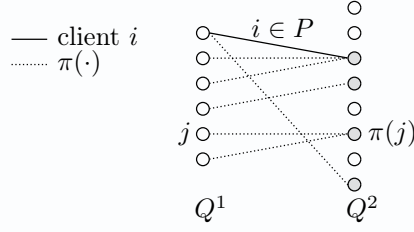
Intuition. We see that a natural rounding algorithm is the following: for an $i \in P$, if its closest facility in Q^1 is opened while its closest facility in Q^2 is not opened, we may just direct i to the opened one in Q^1 , same for the other case. Now, if both facilities are opened, then we direct i to the facility in F^1 with probability a , while to the facility in F^2 with probability $1 - a = b$.

Remark. The problem of the above algorithm is that we don't have control about the total number of the final open facilities: it can be the case that at the end we open every facility in Q^2 , which is k^2 , not k . So we need to sometimes direct i to other facilities (in Q^1) that is not closest to which.

For $j \in Q^1$, let $\pi(j)$ be the closest facility in Q^2 to j , and let Q^* be the image of such a map π , i.e., $Q^* = \{j' \in Q^2 : j' = \pi(j) \text{ for some } j \in Q^1\}$.

Note. We may assume $|Q^*| = k^1$.

Proof. Clearly, $|Q^*| \leq k^1$. And if $|Q^*| < k^1$, we add arbitrary centers so that $|Q^*| = k^1$.



For example, the initial image size above is only 4, we need to add 2 more arbitrary centers into Q^* . ⊛

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

Algorithm 2.18: k -Median – 2-Bipoint Rounding

Data: A set of clients $P \subseteq X$, a set of (possible) facilities $Q \subseteq X$, $a \in (0, 1)$, $\epsilon \in (0, 1)$, $k \in \mathbb{N}$

Result: A set of opened facilities $Q' \subseteq Q$ with $|Q'| = k$

```

1  $(Q^1, Q^2) \leftarrow \text{binary-search}(P, Q, \epsilon)$  // achieve  $|f^1 - f^2| \leq \epsilon \text{OPT} / n$ 
2
3  $Q' \leftarrow \emptyset$ ,  $k^1 \leftarrow |Q^1|$ ,  $k^2 \leftarrow |Q^2|$ ,  $Q^* \leftarrow \{j' \in Q^2 : j' = \pi(j) \text{ for some } j \in Q^1\}$ 
4
5 for  $j \in Q^1$  do
6   if  $\text{rand}(0, 1) \leq a$  then // open  $Q^1$  w.p.  $a$ 
7      $Q' \leftarrow Q' \cup \{j\}$ 
8   else // open  $Q^*$  w.p.  $1 - a$ 
9      $Q' \leftarrow Q' \cup \{\pi(j)\}$ 
10
11 // still need to open  $k - k^1$  more
12  $Q' \leftarrow Q' \cup \{(k - k^1) \text{ random } j \in Q^2 \setminus Q^*\}$ 
13 return  $Q'$ 

```

Remark. Algorithm 2.18 is a randomized algorithm which will **always** open k facilities. The randomness comes from the cost, i.e., we can analyze its cost in expectation.

Intuition. Algorithm 2.18 is kind of *mimicking* what we want, since

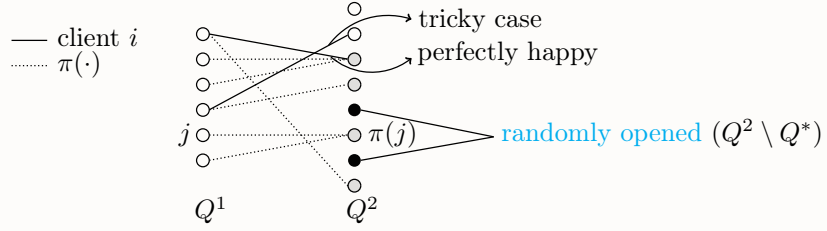
- $j \in Q^1$, $\Pr(j \text{ open}) = a$
- $j \in Q^*$, $\Pr(j \text{ open}) = 1 - a = b$
- $j \in Q^2 \setminus Q^*$, $\Pr(j \text{ open}) = \frac{k - k^1}{k^2 - k^1} = b$

Theorem 2.6.1. Algorithm 2.18 is a 2-bipoint algorithm (in expectation).

Proof. Let's analyze a bit careful. Fixing an $i \in P$, and denote its closest facility in Q^1 as j^1 , and the closest facility in Q^2 as j^2 . If j^1 is not opened, then we know $\pi(j^1)$ is opened for sure in line 8. We see that

- If j^2 is in Q^* , then we know i will be direct to either j^1 or j^2 in line 5, i.e., i is perfectly happy since it can go to one of the closest facility.
- The tricky case is when j^2 is not in Q^* .
 - If j^1 is opened, i can still go to j^1 without problem.
 - If j^1 is also not opened, we know that $\pi(j^1)$ will be opened in line 8. In this worst case,

we just direct i to $\pi(j^1)$ and the distance will be $i \rightarrow j^1 \rightarrow \pi(j^1)$, which is bounded by $d_1 + d(j^1, \pi(j^1))$. But observe that $d(j^1, \pi(j^1)) \leq d_1 + d_2$, so we have $2d_1 + d_2$.



In all, we have the following.^a

	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^b

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

and we now have two cases.

- If $b \geq a$, then $b \geq 1/2$, $M = 0$ and

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

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

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

This shows [Algorithm 2.18](#) is a 2-bipoint algorithm in expectation, proving the result. ■

^aThis is a slightly worse result since we force i to go to j^2 if j^2 is opened, but actually, i can go to j^1 if j^1 is opened too with shorter distance. But this still gives us a good enough bound.

^bSince the final case is always worse than the second case, it is legal to assume that the second case has the minimum probability and the final has the maximum for the expectation bound to hold.

Remark (SOTA). The SOTA result specifically for [Problem 2.6.1](#) is summarized as follows.

Primal-Dual 3-LMP $\xrightarrow{\text{Conversion}}$ 3-approximation

Greedy 2-LMP \longrightarrow 2-bipoint rounding \longrightarrow 4-approximation

Dual Fitting [[Coh+22](#)] 1.9...9-LMP \longrightarrow 1.3...3-bipoint rounding^a \longrightarrow 2.67-approximation

But we'll see that by changing the problem a bit, like consider squaring the distance in the objective of [Problem 2.6.1](#) (which is the k -mean problem), we can get 9-approximation by Primal-Dual, while the lower path doesn't tell us anything, which is so fragile.

^aThis will return $k + c$ centers, where c is an absolute constant. There's a way to transform this solution back to k centers without losing any performance.

Note (Derandomized). It's possible to derandomized [Algorithm 2.18](#).

Lecture 8: Local Search for k -Median

We'll now see a completely different algorithm which solve [Problem 2.6.1](#) with $(3 + \epsilon)$ -approximation ratio by local search. 26 Sep. 10:30

2.6.3 Local Search

The idea is to iteratively improve the current solution. We first see the algorithm.

Algorithm 2.19: k -Median – Local Search

Data: A set of clients $P \subseteq X$, a set of (possible) facilities $Q \subseteq X$, $k \in \mathbb{N}$, width w

Result: A set of opened facilities $Q' \subseteq Q$ with $|Q'| = k$

```

1  $Q' \leftarrow$  arbitrary  $k$  centers in  $Q$ 
2 while  $\exists Q''$  s.t.  $|Q''| = k$  and  $\text{cost}(Q'') < \text{cost}(Q')$  and  $|Q' \Delta Q''| \leq w^a$  do
3    $Q' \leftarrow Q''$ 
4 return  $Q'$ 
  
```

^aThe symmetric difference $A \Delta B$ is defined as $A \Delta B := (A \setminus B) \cup (B \setminus A)$.

Remark (Runtime). In [line 2](#), each iteration in [Algorithm 2.19](#) takes $(n + m)^{O(w)}$ time for $n := |P|$ and $m := |Q|$. But we have no control of how many iterations [Algorithm 2.19](#) might take since we might decrease the cost by a little each time hence we might fall into exponentially many updates. To solve this, we can ask for

$$\text{cost}(Q'') < (1 - \epsilon) \text{cost}(Q')$$

instead to make sure we decrease a reasonable amount each time, which guarantees that we can bound the number of iterations by

$$\log_{\frac{1}{1-\epsilon}} \left(\frac{\text{cost}(\text{starting } Q')}{\text{OPT}} \right).$$

Analysis

Firstly, note that for any solution Q' output from [Algorithm 2.19](#), we have that there exists no Q'' such that $|Q' \Delta Q''| \leq w$, $|Q''| = k$ and $\text{cost}(Q'') < \text{cost}(Q')$.

Note (Local optimum). We say this Q' is a *local optimum*.

Let $Q^* \subseteq Q$ be the optimal solution, and without loss of generality (by duplicating facilities), assume $Q' \cap Q^* = \emptyset$. We define something called [swap](#).

Notation (Swap). A *swap* $S \subseteq Q' \cup Q^*$ satisfies $|S \cap Q'| = |S \cap Q^*| \leq w/2$.

Note. From local optimality of Q' , for any [swap](#) S , $\text{cost}(Q') \leq \text{cost}(Q' \Delta S)$.

Now, consider constructing [swaps](#) S_1, \dots, S_t with weights $p_1, \dots, p_t \in \mathbb{R}^+$ such that $\text{cost}(Q') \leq \text{cost}(Q' \Delta S_i)$ for all i , we have

$$\sum_{i=1}^t p_i \cdot (\text{cost}(Q') - \text{cost}(Q' \Delta S_i)) \leq 0. \quad (2.4)$$

Our goal is to show that [Equation 2.4](#) implies $\text{cost}(Q') \leq \alpha \cdot \text{cost}(Q^*)$ for some $\alpha \in \mathbb{R}^+$. To do this, we require the set of [swaps](#) to have the following properties.

- (a) For all $j \in Q^*$, $\sum_{S_i \ni j} p_i = 1$, and let $p' := \max_{j \in Q'} \sum_{S_i \ni j} p_i$.
- (b) For all $j \in Q^*$, let $\pi(j) \in Q'$ be the facility closest to j . Then if S_i contains $j \in Q'$, $\pi^{-1}(j) \subseteq S_i$.⁴

⁴In particular, if $|\pi^{-1}(j)| > w/2$, then j is not contained in any [swap](#).

The existence of such **swaps** family is ensured by the following lemma.

Lemma 2.6.1. There exists a family of **swaps** S_1, \dots, S_t with weights p_1, \dots, p_t such that $\forall j \in Q^*$, $\sum_{S_i \ni j} p_i = 1$ and if $j \in S_i$, $\pi^{-1}(j) \subseteq S_i$ with $p' = \max_{j \in Q^*} \sum_{S_i \ni j} p_i = 1 + 2/w$.

Proof. For all $j \in Q'$, we call j

- *big*: if $|\pi^{-1}(j)| > w/2$.
- *small*: if $|\pi^{-1}(j)| \in [1, w/2]$.
- *lonely*: if $|\pi^{-1}(j)| = 0$.

Then for each small or big j , we create a *group* G_j that contains $\pi^{-1}(j)$, j and $|\pi^{-1}(j)| - 1$ lonely facilities (denote as $R_j \subseteq Q'$). We see that $|G_j| = 2|\pi^{-1}(j)|$, and we can ensure each lonely facility belongs to exactly 1 group, i.e., $\exists G_1, \dots, G_r$ such that each facility belongs to exactly 1 group. It's now clear that how we should create **swaps** and their corresponding weight:

- (a) For small j , let G_j be a **swap** with weight 1.
- (b) For big j , let $w' := |\pi^{-1}(j)|$, then for any $S \subseteq \pi^{-1}(j)$ and $T \subseteq R_j$ with $|S| = |T| = w/2$, we let $(S \cup T)$ be a **swap** with weight $1 / \left(\binom{w'-1}{w/2-1} \cdot \binom{w'-1}{w/2} \right)$.^a

Since for every $j^* \in Q^*$, there is only one group containing j^* , to verify $\sum_{S_i \ni j^*} p_i = 1$, we see that

- (a) j^* is containing in G_j for j small: In this case, we have one **swap**, i.e., G_j itself with weight 1.
- (b) j^* is containing in G_j for j big: In this case, since every such **swap** created inside G_j contains j^* and has uniform weight, it sums up to 1.

Finally, we want to show that $p' = \max_{j \in Q'} \sum_{S_i \ni j} p_i = 1 + 2/w$. But this is also easy since given $j \in Q'$, the summation is inside G_j , and in particular, S_i is inside G_j as well. Then

- (a) j is small: Only swap is G_j itself with weight 1.
- (b) j is big: j can't even be in one **swap**, hence the sum is 0.
- (c) j is lonely: In this case, we have

$$\sum_{S_i \ni j} p_i = \frac{1}{\binom{w'-1}{w/2-1} \binom{w'-1}{w/2}} \cdot \binom{w'}{w/2} \binom{w'-2}{w/2-1} = \frac{w'}{w/2} \frac{w/2}{w'-1} = \frac{w'}{w'-1} \leq 1 + \frac{2}{w}.$$

Taking the maximum, we have $p' = 1 + 2/w$ as desired. ■

^aNotice that since j is big, so j can't be in any **swap**, so we have only $w' - 1$ to choose from.

With [Lemma 2.6.1](#), we're ready to prove the following.

Theorem 2.6.2. [Algorithm 2.19](#) is a $(3 + \epsilon)$ -approximation algorithm for arbitrary small $\epsilon > 0$.

Proof. Fix $i \in P$, we analyze how it contributes to the left-hand side of [Equation 2.4](#). Let $j' \in Q'$ and $j^* \in Q^*$ be facilities closest to i , and $d'_i := d(i, j')$, $d_i^* = d(i, j^*)$, then for every S_ℓ , we have

- (a) $S_\ell \ni j^*$. Then contribution of i to $\text{cost}(Q') - \text{cost}(S_\ell \triangle Q')$ is at least $d'_i - d_i^*$.^a
- (b) $S_\ell \ni j'$. By the second property, either
 - $\pi(j^*) \in S_\ell$: this implies $j^* \in S_\ell$, which falls back to the first case.
 - $\pi(j^*) \notin S_\ell$: the contribution of i to $\text{cost}(Q') - \text{cost}(S_\ell \triangle Q')$ is at least $d'_i - (2d_i^* + d'_i) = -2d_i^*$.^b
- (c) Otherwise, contribution of i to $\text{cost}(Q') - \text{cost}(S_\ell \triangle Q')$ is at least 0.^c

In all, we see that the first case has total weight 1 from the first property, while (b)–(a) has total weight $\leq p'$, hence Equation 2.4 implies

$$\sum_{i \in P} [(d'_i - d_i^*) \cdot 1 - (2d_i^*) \cdot p'] \leq \sum_{\ell=1}^t p_\ell (\text{cost}(Q') - \text{cost}(S_\ell \triangle Q')) \leq 0,$$

which is equivalent to say

$$\text{cost}(Q') - (1 + 2p') \text{OPT} \leq 0,$$

so we get a $(1 + 2p')$ -approximation ratio. Furthermore, From Lemma 2.6.1, we have $p' = 1 + 2/w$, hence we can achieve $1 + 2(1 + 2/w) = 3 + 4/w$ -approximation ratio. Given $\epsilon > 0$, by setting $w := 4/\epsilon$, we're done. ■

a_i can go to j^* .
 b_i can go to $\pi(j^*)$.
 c_i can stay with j' .

Lecture 9: Euclidean k -Median

2.7 Euclidean k -Median

28 Sep. 10:30

If we now consider the metric space to be exactly $(\mathbb{R}^\ell, \|\cdot\|_2)$, we get some advantages from the structure of Euclidean metric.

Intuition. The problematic part is the old approach for k -median problem is when i contributing to too many facilities at once. But we'll see that this can't happen if we're considering Euclidean metric, which has some inherent geometric limitation in terms of volume.

Now, let's see the problem formulation.

Problem 2.7.1 (Euclidean k -median). Given a metric space $(X, d) = (\mathbb{R}^\ell, \|\cdot\|_2)$ and $P, Q \subseteq X$ with $k \in \mathbb{N}$, find $Q' \subseteq Q$ with $|Q'| = k$ which minimizes $\sum_{i \in P} \min_{j \in Q'} d(i, j)$.

It's natural to ask that whether we can solve Euclidean k -median like how we solve facility location and k -median. The answer is yes, and in particular, we're going to modify ?? for facility location to get an α -LMP approximation with $\alpha < 3$, which essentially implies an α -approximation algorithm for k -median using Euclidean metric.

2.7.1 Euclidean Facility Location

Formally, we define the following problem.

Problem 2.7.2 (Euclidean facility location). Given a metric space $(X, d) = (\mathbb{R}^\ell, \|\cdot\|_2)$ 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 $Q' \subseteq Q$ such that it minimizes $\sum_{i \in P} \min_{j \in Q'} d(i, j) + f |Q'|$.

As previously seen. Recall the dual of facility location is

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

and the ?? uses primal-dual method, where we interpret α_i is the time that i is connected.

Let t_j be the time that j is open in ??, and the only thing we change is the phase two, i.e., how we trim down the solution. We now see the algorithm, which essentially achieves $\rho := (1 + \delta)$ -LMP approximation for $\delta := \sqrt{8/3} \approx 1.633 \dots$

Algorithm 2.20: Euclidean 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 $Q' \subseteq Q$

```

1  $S \leftarrow \emptyset, Q' \leftarrow \emptyset, \alpha \leftarrow 0$  //  $S$ :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  $j \in Q \setminus Q'$  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  $j \in Q'$ 
9       break
10     $Q' \leftarrow \{\text{tight facilities}\}$  // Update  $Q'$ 
11     $S \leftarrow \{\text{clients connected to } Q'\}$  // Update  $S$ 
12
13 // Trim down  $Q'$ 
14  $G = (Q', E := \{(j, j') : \exists i \in P \text{ such that } d(j, j') \leq \delta \cdot \min(t_j, t_{j'}), j, j' \in Q'\})$ 
15 Compute  $Q''$  s.t.  $\forall j \in Q'$ , either  $j \in Q''$  or  $\exists j' \in Q''$  s.t.  $(j, j') \in E$  // max independent set
16 return  $Q''$ 

```

Analysis

As before, let $w(i) \in Q'$ for all i such that $\alpha \geq t_{w(i)}$, i.e., $w(i)$ is the connected witness of i .

Remark. We have the following.

- (a) α is dual-feasible.
- (b) If $\beta_{ij} > 0$, then $\alpha_i \leq t_j$.
- (c) For all i , $\exists w(i) \in Q'$ such that $\alpha_i \geq t_{w(i)}$.

We can do the analysis similarly. Fix a client $i \in P$, then observe that given $S = Q'' \cap \{j : \beta_{ij} > 0\}$, if $\delta = 2$, then $|S| \leq 1$.⁵ We see that

- (a) If $|S| = 1, S = \{j\}$. We see that $\text{conn}(i) \leq d(i, j)$ and $\text{open}(i) = \alpha_i - d(i, j)$, so

$$\text{conn}(i) + \text{open}(i) \leq d(i, j) + (\alpha_i - d(i, j)) \leq \alpha_i.$$

- (b) If $|S| = 0$, then $\text{open}(i) = 0$ and either $w(i) \in Q''$, or $j' \in Q''$ such that $(w(i), j') \in E$. In any case, $\text{conn}(i) \leq d(i, j') \leq d(i, w(i)) + d(w(i), j')$, hence

$$\text{conn}(i) + \text{open}(i) \leq d(i, j') \leq \alpha_i + \delta t_{w(i)} \leq (1 + \delta)\alpha_i.$$

Generally, our goal is to prove that for all i ,

$$\frac{\text{conn}(i)}{\rho} + \text{open}(i) \leq \alpha_i, \tag{2.5}$$

which implies

$$\frac{\text{conn}}{\rho} + |Q''|f \leq \sum_i \alpha_i,$$

i.e., we get a ρ -LMP approximation algorithm.

⁵Since if both β_j and $\beta_{j'}$ is greater than 0, then $d(j, j') \leq 2\alpha_i \leq 2\min(t_j, t_{j'})$. This means j and j' will have an edge but from the property of max independent set, one of them will not be included.

Note. Specifically, Equation 2.5 is equivalent to

$$\frac{\min_{j \in S} d(i, j)}{\rho} + \sum_{j \in S} (\alpha_i - d(i, j)) \leq \alpha_i.$$

In the case of $\delta = 2$, we see that we can set $\rho := 1 + \delta = 3$. We see that we get the exactly 3-LMP approximation for $\delta = 2$ case! Notice that in this case, since $|S| \leq 1$ as we noted, Algorithm 2.20 and ?? are equivalent.

Now, we'll see how can we get advantages by further restricting δ , which utilizes the following.

Remark (k -means magic formulas). There are two extremely useful tricks call *k-means magic formulas* for Euclidean metric related problems. Let $i' = \sum_{j \in S} j / |S|$. Then

$$\begin{aligned} \sum_{j \in S} \|j - i\|^2 &= \sum_{j \in S} \langle j - i + i' - i', j - i + i' - i' \rangle \\ &= \sum_{j \in S} \left(\|j - i\|^2 + \|i' - i\|^2 + 2 \langle j - i', i' - i \rangle \right) = \sum_{j \in S} \|j - i'\|^2 + |S| \|i' - i\|^2. \end{aligned}$$

Also,

$$\begin{aligned} \sum_{j, j' \in S} \|j - j'\|^2 &= \sum_{j, j' \in S} \langle j - j' + i' - i', j - j' + i' - i' \rangle \\ &= \sum_{j, j' \in S} \left(\|j - i'\|^2 + \|j' - i'\|^2 + 2 \langle j - i', i' - j' \rangle \right) = 2|S| \cdot \sum_{j \in S} \|j - i'\|^2. \end{aligned}$$

One can actually show that i' (i.e., the geometric mean) is the optimal solution for k -means, and if we choose i rather than i' to be the center, the deviation from OPT is exactly $|S| \|i' - i\|^2$. Nevertheless, we have the following.

Lemma 2.7.1. For $\delta := \sqrt{8/3}$ and $S = Q'' \cap \{j : \beta_{ij} > 0\}$, $|S| \leq 3$.

Proof. From the k -means magic formulas, we have

$$|S| \alpha_i^2 \geq \sum_{j \in S} \|j - i\|^2 \geq \frac{1}{2|S|} \sum_{j, j' \in S} \|j - j'\|^2 > \frac{(s-1)\delta^2 \alpha_i^2}{2},$$

where the last inequality follows from $\|j - j'\| > \delta \cdot \min(t_j, t_{j'}) \geq \delta \cdot \alpha_i$. Then, we have

$$|S| \alpha_i^2 > \frac{(s-1)\delta^2 \alpha_i^2}{2} \Rightarrow |S| \left(\frac{\delta^2}{2} - 1 \right) < \frac{\delta^2}{2} \Rightarrow |S| < \frac{\delta^2}{\delta^2 - 2} = 4$$

by plugging in $\delta = \sqrt{8/3}$, hence $|S| \leq 3$ by integrality. \blacksquare

From Lemma 2.7.1, we see that we already handle the case that $|S| = 0$ and $|S| = 1$, so the only cases left are $|S| = 2$ and $|S| = 3$. And by doing so, we obtain the following theorem.

Theorem 2.7.1. Algorithm 2.20 is a $(1 + \sqrt{8/3})$ -LMP approximation algorithm.

Proof. As said, from Lemma 2.7.1, we only need to consider the case that $|S| = 2$ and $|S| = 3$. If $|S| = 2$, let $S = \{j_1, j_2\}$, then $(\alpha_i - d(i, j_1)) + (\alpha_i - d(i, j_2)) \leq (2 - \delta)\alpha_i$. Since $\text{conn}(i) \leq \alpha_i$,

$$\frac{d(i, j^*)}{\rho} + \sum_{j \in S} (\alpha_i - d(i, j)) \leq \alpha_i \left(\frac{1}{\rho} + 2 - \delta \right) \leq \alpha_i, \quad (2.6)$$

where the last inequality follows from $1/\rho + 2 - \delta \leq 1 \Leftrightarrow 1/\rho \leq \delta - 1$, which is satisfied by $\rho := 1 + \delta = 1 + \sqrt{8/3}$.

If $|S| = 3$, let $S = \{j_1, j_2, j_3\}$. Now, instead of looking at a more complicated geometric structure and try to optimize it, we simply add Equation 2.6 three times for (j_1, j_2) , (j_2, j_3) and (j_1, j_3) , we have $2 \sum_{j \in S} (\alpha_i - d(i, j)) \leq 3(2 - \delta)\alpha_i$ hence

$$\frac{d(i, j^*)}{\rho} + \sum_{j \in S} (\alpha_i - d(i, j)) \leq \alpha_i \left(\frac{1}{\rho} + \frac{3(2 - \delta)}{2} \right) \leq \alpha_i$$

since $1/\rho + 3(2 - \delta)/2 \leq 1 \Leftrightarrow 1/\rho \leq (3\delta - 4)/2$, which is satisfied by $\rho := 1 + \delta = 1 + \sqrt{8/3}$ ■

Remark (SOTA). Compare general metric Problem 2.6.1 and Euclidean metric Problem 2.7.1, we have the following.

2.41-LMP^a $\xrightarrow{\text{Conversion}}$ 2.41-approximation

Euclidean Primal-Dual 2.63-LMP $\xrightarrow{\text{Conversion}}$ 2.63-approximation

Primal-Dual 3-LMP $\xrightarrow{\text{Conversion}}$ 3-approximation

Dual Fitting [Coh+22] 1.9...9-LMP \longrightarrow 1.3...3-bipoint rounding \longrightarrow 2.67-approximation

Local Search 3-LMP \longrightarrow 2-bipoint rounding \longrightarrow 4-approximation

Noticeably, 2.41-LMP approximation is $1 + \sqrt{2}$, which is exactly the threshold behavior in Euclidean metric we're building our intuition upon.

^a2.40 is the SOTA.

Note. We assume that ℓ is large throughout. If it's not the case, then actually for all $\epsilon > 0$, there exists a $(1 + \epsilon)$ -approximation algorithm with running time $2^{2^{O(\ell)}} \cdot \text{poly}(n)$. Hence, if ℓ is small (or constant), we can use this algorithm, otherwise, what we have discussed is better.

Chapter 3

Traveling Salesman Problem

Lecture 10: Spanning Tree

Instead of discussing general network design problems, we focus on traveling salesman problem specifically. And turns out that although this is a good old problem in TCS, but still, lots of improvement is done in the past decade. Turns out, most of the improvement is based on the understanding of spanning tree, specifically, how to sample a good enough random [spanning tree](#).

3 Oct. 10:30

3.1 Spanning Tree

We first look at the definition of a [spanning tree](#).

Definition 3.1.1 (Spanning tree). A *spanning tree* T of a connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ is an induced subgraph of \mathcal{G} which spans \mathcal{G} , i.e., $V(T) = \mathcal{V}$ and $E(T) \subseteq \mathcal{E}$.

Remark. A [spanning tree](#) of a connected graph \mathcal{G} can also be defined as a maximal set of edges of \mathcal{G} that contains no cycle, or as a minimal set of edges that connect all vertices.

Then, we're interested in the following problem.

Problem 3.1.1 (Minimum spanning tree). Given a connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ and an edge-weight function $w: \mathcal{E} \rightarrow \mathbb{R}^+$, find a [spanning tree](#) T which minimizes $w(T)$.

There are lots of different algorithms which solve [Problem 3.1.1](#), e.g., [Prim's algorithm](#), [Kruskal's Algorithm](#), etc. in undergraduate algorithm courses. But turns out that by looking at the LP formulation of this problem, we get some non-trivial result.

3.1.1 Spanning Tree Polytope

Denote the variables as $\{x_e\}_{e \in \mathcal{E}}$, where we interpret $x_e = 1$ if e is in the final [spanning tree](#), otherwise if it's 0, then e is not in the final [spanning tree](#).

One natural formulation is

$$\begin{aligned} \min \quad & \sum_{e \in \mathcal{E}} x_e w(e) \\ & \sum_{e \in \partial S} x_e \geq 1 \quad \forall S \subseteq \mathcal{V} \\ & x \geq 0, \end{aligned}$$

where the second constraint is trying to model that for every cut set $S \subseteq \mathcal{V}$, our [spanning tree](#) need to include at least one edge from the boundary, i.e., ∂S .

Notation. If $S \subseteq \mathcal{V}$, then we denote $\partial S = E(S, \bar{S})$ be the edges between S and \bar{S} .

But turns out that this formulation will give us an **integrality gap** of 2, since for a cycle graph, just by choosing half of the edges, i.e., $x_e = 1/2$ for all $e \in \mathcal{E}$, the constraints are satisfied while we know we need to include all but one edge to form a valid **spanning tree**.

Remark. There are ways to strengthen the second constraints by looking at **directed spanning trees** rather than the usual undirected ones to give us an LP which solves **Problem 3.1.1** exactly.

We see that the problems arise from the fact that there are not enough edges to span \mathcal{G} , so we now require it explicitly in our LP formulation. Furthermore, to ensure there are no cycles, for any $S \subseteq \mathcal{V}$, we again make sure that the total edges we have is less than $|S| - 1$. Then, we have the following **spanning tree** polytope.

$$\begin{aligned} \min \quad & \sum_{e \in \mathcal{E}} x_e w(e) \\ & \sum_{e \in \mathcal{E}} x_e = n - 1 \\ & \sum_{e \in E(S)} x_e \leq |S| - 1 \quad \forall \emptyset \neq S \subsetneq \mathcal{V} \\ & x \geq 0 \end{aligned} \tag{3.1}$$

where $E(S)$ denotes the set of edges inside S , i.e.,

$$E(S) := \{e = (u, v) \in \mathcal{E} : u, v \in \mathcal{V}\}.$$

This is not solvable just by throwing this into an LP solver since there are exponentially many constraints! Regardless, we note the following.

Remark (Separation oracle). Given a linear program (P) with $x \in \mathbb{R}^n$ as variables, a *separation oracle* is an algorithm which outputs

- Yes if x is feasible.
- No with *the violating constraint* if x is not feasible.

And if we have a polynomial time separation oracle, we can solve any LP in polynomial time by using the **ellipsoid algorithm**.

Now, we just state that there's a **separation oracle** for the above LP, so we can solve it in polynomial time and get a fractional solution $\{x_e\}_{e \in \mathcal{E}}$. So our next task is to round it into an integral one.

3.1.2 Pipeage Rounding

The reason we call **Equation 3.1** the *polytope* is that there's a way to transform the any optimal (potentially fractional) solution to this LP can be transformed to an integral one while maintaining the objective value. This can be done via the so-called *pipeage rounding* as we'll now see.

Notation (Tight). The set $S \subseteq \mathcal{V}$ is *tight* if $\sum_{e \in E(S, \bar{S})} x_e = |S| - 1$.

Lemma 3.1.1 (Uncrossing). If S and T are **tight** with $S \cap T \neq \emptyset$, both $S \cup T$ and $S \cap T$ are **tight**.

Proof. Observe that since S and T are **tight** and $S \cup T$ and $S \cap T$ are cuts as well (hence satisfy the constraints),

$$\begin{aligned} (|S| - 1) + (|T| - 1) &= \sum_{e \in E(S)} x_e + \sum_{e \in E(T)} x_e \\ &\leq \sum_{e \in E(S \cup T)} x_e + \sum_{e \in E(S \cap T)} x_e \leq (|S \cup T| - 1) + (|S \cap T| - 1), \end{aligned}$$

with the fact that $|S| + |T| = |S \cup T| + |S \cap T|$,^a hence everything is equal. ■

^aConsider every possible edges between $S \setminus T$, $T \setminus S$, $S \cap T$ and $\overline{S \cup T}$.

Finally, we call a **tight** T *integral* if and only if for all $e \in E(T)$, $x_e \in \{0, 1\}$; and a **tight** T *fractional* if there exists $e \neq f \in E(T)$ such that x_e and x_f are fractional.¹ We first see the deterministic rounding algorithm.

Algorithm 3.1: Minimum Spanning Tree – Pipage-Rounding

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, edge weight $w: \mathcal{E} \rightarrow \mathbb{R}^+$, solution x of Equation 3.1^a

Result: A minimum **spanning tree** T

```

1 while  $x \notin \mathbb{N}^m$  do                                     // not integral
2    $T \leftarrow$  minimal tight fraction set                 // inclusion-wise minimalb
3    $f, g \leftarrow$  fractional edges                       //  $f, g \in E(T)$ 
4   if  $w(f) > w(g)$  then                                   // ensure  $w(f) \leq w(g)$ 
5     swap( $f, g$ )
6   while increase  $x_f$  and decrease  $x_g$  by a unit do     // by solving Equation 3.2
7     if  $x_f$  or  $x_g$  becomes integral then
8       break
9     else if  $\exists T' \subsetneq T$  is tight then
10      break
11  $T \leftarrow$  Subgraph( $\mathcal{G}, x$ )                             // construct a spanning tree
12 return  $T$ 
```

^aBy using **separation oracle**.

^bi.e., $\nexists T' \subsetneq T$ **tight** fractional set.

Remark (Implementation detail). There are two non-trivial steps in Algorithm 3.1.

- **line 6:** This continuous process is done by taking δ from solving the following LP as the total unit we should increase/decrease:

$$\begin{aligned}
 \max \quad & \delta \\
 \text{s.t.} \quad & y = x + \delta e_f - \delta e_g \\
 & \sum_{e \in E(S)} y - e \leq |S| - 1 \quad \forall S \subseteq \mathcal{V} \\
 & 0 \leq y \leq 1,
 \end{aligned} \tag{3.2}$$

where e_i is the unit vector with 1 at entry i . Again, this is in the similar form as Equation 3.1, and there's a **separation oracle** which solves this LP in polynomial-time.

- **line 2:** Start from the whole vertex set \mathcal{V} , and we simply look at f which is none-integral edge and ask can we increase it or not, i.e., we ask the **separation oracle** for Equation 3.2, and if there's a smaller **tight** fraction set inside T , $\delta > 0$ strictly, and we just keep searching in this way. We'll see what does this mean exactly in Lemma 3.1.2.

Our goal now is to show that during the **pipage rounding**, x remains feasible and $\sum_{e \in \mathcal{E}} x_e w(e)$ will not increase.

We first show that $\sum_{e \in \mathcal{E}} x_e w(e)$ will not increase. This is because from our design, $\sum_{e \in \mathcal{E}} x_e$ remains unchanged, while we increase x_f while decrease x_g for $w(f) \leq w(g)$, hence the total cost for the **spanning tree** decreases.

To show x remains feasible, first note that the non-**tight** sets are handled (captured) in line 9, as for **tight** sets, we have Lemma 3.1.2.

Lemma 3.1.2. All **tight** sets remain **tight** after running line 6.

¹We can equivalently require only one x_e being fractional, but since T is **tight**, there'll another $f \neq e$ such that x_f is fractional as well.

Proof. The only way for a **tight** set becomes over-**tight** is when we increase x_f in **line 6**, an already **tight** set U becomes over-**tight**. But if this is the case and U is violated, then $U \ni f$ and $U \not\ni g$ and U is **tight**, we have $U \cap T$ is **tight** from **Lemma 3.1.1**, contradicting the minimality of T \nmid ■

Remark. From the poof, we can now **find minimal T** by increasing a fractional x_f : if some set U is not violated, then $T \cap U$ is **tight**, so we just keep nesting and get the minimal one.

Now, it remains to show **Algorithm 3.1** terminates in polynomial time.

Lemma 3.1.3. **Algorithm 3.1** in a polynomial time algorithm.

Proof. Observe that

- (a) **line 7** can only happen m times: at most m edges can be fractional at first, and after one becomes integral, it remains integral.
- (b) **line 9** can only happen n times: at most n nodes can be in T at first, and when **line 9** is triggered, the size of T decreases by at least 1 and never goes up.

In all, we see that **Algorithm 3.1** is a polynomial time algorithm. ■

Note. Notice that in **line 9**, we require $T' \subsetneq T$, and if it's triggered, in the next iteration when choosing T in **line 2**, we'll need to choose a strictly smaller T compare to the last iteration^a in order to make **Lemma 3.1.3** valid.

^aThis is guaranteed by **Lemma 3.1.2** since we know the only we change is x_f and x_g , and if some new T' can become **tight**, it has non-empty intersection with T and hence as the **remark**, we can find such a T' .

We see that this implies the following.

Theorem 3.1.1. **Algorithm 3.1** solves **Problem 3.1.1** exactly in polynomial time.

Proof. Firstly, **Algorithm 3.1** is a polynomial time algorithm from **Lemma 3.1.3**. Also, since **Equation 3.1** is an LP-relaxation of **Problem 3.1.1** while we know that ■

And indeed, we have a randomized version of **Algorithm 3.1**.

Algorithm 3.2: Minimum Spanning Tree – Randomized Pipage-Rounding

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, edge weight $w: \mathcal{E} \rightarrow \mathbb{R}^+$, solution x of **Equation 3.1**^a

Result: A minimum **spanning tree** T

```

1 while  $x \notin \mathbb{N}^m$  do                                     // not integral
2    $T \leftarrow$  minimal tight fraction set                // inclusion-wise minimalb
3    $f, g \leftarrow$  fractional edges                        //  $f, g \in E(T)$ 
4   if  $w(f) > w(g)$  then                                    // ensure  $w(f) \leq w(g)$ 
5     swap( $f, g$ )
6    $a \leftarrow \max_a x_f \leftarrow x_f + a, x_g \leftarrow x_g - a$  remain feasible //  $a > 0$ 
7    $b \leftarrow \max_b x_f \leftarrow x_f - b, x_g \leftarrow x_g + b$  remain feasible //  $b > 0$ 
8   if  $\text{rand}((0, 1)) < \frac{b}{a+b}$  then                      // w.p.  $\frac{b}{a+b}$ 
9      $x_f \leftarrow x_f + a, x_g \leftarrow x_g - a$ 
10  else                                                    // w.p.  $\frac{a}{a+b}$ 
11     $x_f \leftarrow x_f - b, x_g \leftarrow x_g + b$ 
12  $T \leftarrow \text{Subgraph}(\mathcal{G}, x)$                         // construct a spanning tree
13 return  $T$ 

```

^aAgain, by using **separation oracle**.

^bi.e., $\nexists T' \subsetneq T$ **tight** fractional set.

As in the **deterministic version**, the same proof can show that x is feasible, and the number of iteration will be less than $m \cdot n$, hence it's a polynomial time algorithm. Remarkably, we have the following.

Theorem 3.1.2. Algorithm 3.2 solves Problem 3.1.1 exactly.

Proof. To show that the cost is good enough, note that in one iteration, $\mathbb{E}[x^{\text{end}}] = x^{\text{start}}$, then

$$\mathbb{E}[x^{\text{final}}] = x^{\text{LP}},$$

hence any possible x^{final} satisfies

$$\sum_{e \in \mathcal{E}} x_e^{\text{final}} w(e) = \sum_{e \in \mathcal{E}} x_e^{\text{LP}} w(e),$$

hence we get a 1-approximation algorithm, i.e., Algorithm 3.2 solves Problem 3.1.1 exactly. ■

From Algorithm 3.2, x^{final} can be interpreted as the distribution of spanning trees, i.e., we have

$$\mathbb{E}[x^{\text{final}}] = x^{\text{LP}} \Leftrightarrow \forall e \in \mathcal{E}, \Pr(e \in T) = x_e^{\text{LP}},$$

where the probability depends on the randomness introduced in Algorithm 3.2, i.e., x_e^{final} . So, from now on, when we say we sample a spanning tree from x , what we mean is to construct a spanning tree w.r.t. the solution x to the spanning tree polytope using Algorithm 3.2.

3.2 Negative Correlation

One of the reasons why we're interested in Algorithm 3.2 is because it produces a negative correlated distribution, which leads to a strong concentration behavior, i.e., we have control on what kind of spanning tree we're going to get. Firstly, if x_e^{final} are independent, then

$$\mathbb{E}\left[\prod_{e \in S} x_e^{\text{final}}\right] = \Pr(S \subseteq T) = \prod_{e \in S} \Pr(e \in T) = \prod_{e \in S} x_e^{\text{LP}}.$$

But since we know that x_e^{final} are not independent for sure since they depend on a sequence of steps executed by Algorithm 3.2, it's non-trivial to analyze. We now see the main result in this section.

Theorem 3.2.1 (Negative correlation). For all $S \subseteq \mathcal{E}$,

$$\mathbb{E}\left[\prod_{e \in S} x_e^{\text{final}}\right] = \Pr(S \subseteq T) \leq \prod_{e \in S} \Pr(e \in T) = \prod_{e \in S} x_e^{\text{LP}}.$$

Proof. Let y^i be x after i^{th} iteration maintained by Algorithm 3.2, it's sufficient to show

$$\mathbb{E}\left[\prod_{e \in S} y_e^{i+1} \mid y^i\right] \leq \prod_{e \in S} y_e^i$$

since if this holds, say Algorithm 3.2 runs M iterations in total, then

$$\mathbb{E}\left[\prod_{e \in S} x_e^{\text{final}}\right] = \mathbb{E}\left[\prod_{e \in S} y_e^M\right] = \mathbb{E}\left[\prod_{e \in S} y_e^M \mid y^{M-1}\right] \leq \prod_{e \in S} y_e^{M-1},$$

any by taking expectation again iteratively, we obtain the desired result down to $\prod_{e \in S} y_e^0$. Now, consider that in the i^{th} iteration of Algorithm 3.2, for f, g picked in line 3:

(i) $f, g \notin S$: trivially holds.

(ii) $f \in S, g \notin S$:^a we have $\mathbb{E}[\prod_{e \in S} y_e^{i+1} \mid y^i] = \prod_{e \in S \setminus \{f\}} y_e^i \cdot \mathbb{E}[y_f^{i+1} \mid y^i] = \prod_{e \in S} y_e^i$ where $\mathbb{E}[y_f^{i+1} \mid y^i] = y_f^i$ is the designed from Algorithm 3.2.

(iii) $f, g \in S$. Suffices to compare $\mathbb{E}[y_f^{i+1} \cdot y_g^{i+1} \mid y^i]$ and $y_f^i \cdot y_g^i$, and the goal is to show \leq .

- (a) $\mathbb{E}[(y_f^{i+1} + y_g^{i+1})^2 \mid y^i] = (y_f^i + y_g^i)^2$ since $y_f^{i+1} + y_g^{i+1} = y_f^i + y_g^i$ almost surely.
- (b) $\mathbb{E}[(y_f^{i+1} - y_g^{i+1})^2 \mid y^i] \geq (y_f^i + y_g^i)^2$ since the variance of any random variable is non-negative.

We see that by subtracting them, we have $\mathbb{E}[y_f^{i+1} \cdot y_g^{i+1} \mid y^i] \leq y_f^i \cdot y_g^i$ as desired.

In all cases, the hypothesis for i holds, hence the theorem is proved. \blacksquare

^aAnd also $g \in S$ and $f \notin S$, since they're symmetric.

Lecture 11: Asymmetric TSP

As previously seen. We have shown that given any feasible x , there's a distribution of [spanning tree](#) T such that

5 Oct. 10:30

- (a) For all $e \in \mathcal{E}$, $\Pr(e \in T) = x_e$
- (b) For all $S \subseteq \mathcal{E}$, $\Pr(S \subseteq T) \leq \prod_{e \in S} x_e$ from [Theorem 3.2.1](#).

From these, we can deduce the following.

Theorem 3.2.2. For all $S \subseteq \mathcal{E}$ and $\gamma \geq 1$,

$$\Pr\left(|S \cap T| \geq \gamma \sum_{e \in S} x_e\right) \leq \left(\frac{e}{\gamma}\right)^{\gamma \sum_{e \in S} x_e}.$$

Proof. This follows directly from the same proof of [Chernoff bound](#). Assume we have k random variables $X_1, \dots, X_k \in \{0, 1\}$ with $X = \sum_{i=1}^k X_i$ and $\mu = \mathbb{E}[X]$. Then,

$$\Pr(X \geq \gamma\mu) = \Pr(e^{tX} \geq e^{t\gamma\mu}) \leq \frac{\mathbb{E}\left[\prod_{i=1}^k e^{tX_i}\right]}{e^{t\gamma\mu}}$$

where the inequality follows from Markov's inequality. If X_i are independent, we can move the expectation inside the product, but if we don't, we directly apply [Theorem 3.2.1](#) to get the same result, so we can proceed as usual. \blacksquare

3.3 Asymmetric Traveling Salesman Problem

Now we can talk about the [asymmetric traveling salesman problem](#). Before we state the problem, we first look at one important definition.

Definition 3.3.1 (Tour). Given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, a *tour* (a_0, \dots, a_k) where $a_i \in \mathcal{V}$ satisfies $a_0 = a_k$, $(a_i, a_{i+1}) \in \mathcal{E}$ and visited all the vertices, i.e., $\{a_i\}_{i=0}^k = \mathcal{V}$.

Problem 3.3.1 (Asymmetric TSP). Given a complete bidirected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ and a distance function $d: \mathcal{E} \rightarrow \mathbb{R}^+$ satisfying the *directed* triangle inequality.^a *Asymmetric TSP* asks to find a *tour* (a_0, \dots, a_k) which minimizes $\sum_{i=0}^{k-1} d(a_{i-1}, a_i)$.

^aCompare to the regular triangle inequality, now the order matters, i.e., for all $a, b, c \in \mathcal{V}$, $d(a, c) \leq d(a, b) + d(b, c)$.

Remark. An equivalent (but seemingly more general) formulation of [Problem 3.3.1](#) is to remove the complete graph restriction and also the directed triangle inequality property of d . But they're still

equivalent since given this general problem, we can convert back to [Problem 3.3.1](#) by setting

$$d'(u, v) = \min d(u, v).$$

Note (SOTA). The approximation ratio of [Problem 3.3.1](#) is improved as follows.

$$\begin{array}{ccc} \lg n & \xrightarrow{1982 \sim 2009} & c \lg n \xrightarrow{2009 \sim 2010} O(\log n / \log \log n) \\ & & \downarrow 2010 \sim 2018 \\ & & 5500 \xrightarrow{2018 \sim 2020} 22 \end{array}$$

where in 2009, $c \in (0, 1)$.

3.3.1 Asymmetric TSP Polytope

We now try to solve [Problem 3.3.1](#). The idea is simple, given $T \subseteq \mathcal{E}$ for T being a multiset, we want T to satisfy

- (a) T is connected (in undirected sense)
- (b) T is **Eulerian**: $\deg_T^+(v) = \deg_T^-(v)$ for all $v \in \mathcal{V}$ ²

which allow us to potentially construct a valid **tour** by shortcut some repetitions if there's any. We then have the following LP formulation, which is the so-called **asymmetric TSP polytope**. Denote our variables as $\{x_e\}_{e \in \mathcal{E}}$, then

$$\begin{aligned} \min \quad & \sum_{e \in \mathcal{E}} x_e d(e) \\ \text{s.t.} \quad & \sum_{e \in \partial^+ S} x_e \geq 1 & \forall \emptyset \neq S \subsetneq \mathcal{V} \\ & \sum_{e \in \partial^+ \{v\}} x_e = \sum_{e \in \partial^- \{v\}} x_e = 1 & \forall v \in \mathcal{V} \\ & x \geq 0 \end{aligned} \tag{3.3}$$

where $\partial^+ S := \{(u, v) \in \mathcal{E} \mid u \in S, v \notin S\}$ and vice versa, and $\partial S := \partial^+ S \cup \partial^- S$. Now, to solve this LP, the idea is to maintain **Eulerianity** while gradually being more connected. This can be done via cycle cover LP.

As previously seen. Recall that $C \subseteq \mathcal{E}$ is a cycle cover if c is disjoint union of directed cycles and v is in exactly one cycle.

Now, we have the following.

Lemma 3.3.1. There is a cycle cover C such that $\sum_{e \in C} d(e) \leq \text{OPT}_{\text{LP}}$.

To prove [Lemma 3.3.1](#), we need to have some understanding about the **perfect matching polytope**. This is not that well-known since matching problem can be solved in many ways.

Remark (Perfect matching polytope). Suppose we have an unweighted bipartite graph $\mathcal{G} = (A \sqcup B, \mathcal{E})$ and a weight function $w: \mathcal{E} \rightarrow \mathbb{R}^+$. We want to find a perfecting matching with minimum cost.

²We use \deg^+ to denote the out degree, while \deg^- to denote the in degree.

This can be modeled by the following LP.

$$\begin{aligned} \min \quad & \sum_{e \in \mathcal{E}} x_e w(e) \\ \text{s.t.} \quad & \sum_{e \in E(u, v-u)} x_e = 1 \quad u \in A \sqcup B \\ & x \geq 0. \end{aligned}$$

This LP is exact in the sense that for any feasible x , there exists a perfect matching (integral solution) M such that $\sum_{e \in M} w(e) \leq \sum_{e \in \mathcal{E}} x_e w(e)$.

Now we can prove [Lemma 3.3.1](#).

Proof of Lemma 3.3.1. We simply construct a complete bipartite graph with vertex set $\mathcal{V}_{\text{out}} \sqcup \mathcal{V}_{\text{in}}$ such that the $\mathcal{V}_{\text{out}} = \mathcal{V}_{\text{in}} = \mathcal{V}$ with the edge weight being $x_{(a,b)}$ for a in the left-hand side while b in the right-hand side.

Observe that in B , every vertex has x value exactly 1, hence from [perfect matching polytope](#), we know that there exists a perfect matching M in B with cost $(\sum_{e \in M} w(e))$ less the LP cost $(\sum_{e \in \mathcal{E}} x_e w(e))$, with the fact that M corresponds to a cycle cover in the original graph by considering picking $(a, b) \in M$ the directed edge $(a, b) \in \mathcal{E}$, so we're done. ■

Then, we have the following algorithm.

Algorithm 3.3: [Asymmetric TSP](#) – Cycle Covered

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, distance function $d: \mathcal{E} \rightarrow \mathbb{R}^+$

Result: A [tour](#) T

```

1  $\mathcal{C} \leftarrow$  minimum cycle cover of  $\mathcal{G}$ 
2  $\mathcal{V}' \leftarrow \emptyset$ 
3 for  $C \in \mathcal{C}$  do
4    $x \leftarrow \text{rand}(C)$  // Choose one representative
5    $\mathcal{V}' \leftarrow \mathcal{V}' \cup \{x\}$ 
6  $T \leftarrow \text{ATSP}(\mathcal{G}[\mathcal{V}'], d)$  // tour among representatives
7 return  $T \leftarrow \text{Stitch}(\mathcal{C}, T)$ 
8
9 Stitch( $\mathcal{C}, T$ ):
10   for  $C \in \mathcal{C}$  do
11      $T \leftarrow T \cup C$  // Connects  $T$  with  $C$ 
12   return  $T$ 
```

Theorem 3.3.1. [Algorithm 3.3](#) is a $\lg n$ -approximation algorithm.

Proof. We simply observe that for every recursive call of solving cycle cover LP, since we don't have self-loops, so the number of vertices, \mathcal{V}' , constructed in [Algorithm 3.3](#) will decrease by a factor of 2 since every cycle need at least two vertices, so the total number of recursive calls will be at most $\lg n$. From the fact that in each recursive call, the cost will be at most the cost of the original LP solution for the entire graph from [Lemma 3.3.1](#),^a so by adding the cost up (i.e., stitching the tour together), the total cost will be at most $\lg n \cdot \text{OPT}$, proving the result. ■

^aRecall that we're recursively solving for subgraph of \mathcal{G} .

Remark (Repetition). Observing that in [Algorithm 3.3](#), our construction might not return a valid [Eulerian tour](#). But by triangle inequality, we can always skip some vertices when we encounter already visited vertices, so we're still fine.

3.3.2 Reducing to Thin Tree

Now let's see more sophisticated approach to [Problem 3.3.1](#) where we first make sure T is connected, and try to make it [Eulerian](#) afterwards. One problematic case is that when there's one $S \subseteq \mathcal{V}$ such that T has lots of edges in ∂S , then since we want to ensure $\deg_T^+(v) = \deg_T^-(v)$ for all $v \in \mathcal{V}$, by summing up for all v , we'll need to balance this out by (potentially) adding lots of edges on top of T to make it [Eulerian](#).

Definition 3.3.2 ((α, β) -thin). A tree $T \subseteq \mathcal{E}$ is (α, β) -thin if $\sum_{e \in T} d(e) \leq \alpha \text{OPT}$ and $|T \cap \partial S| \leq \beta \sum_{e \in \partial S} x_e$ for all $\emptyset \neq S \subsetneq \mathcal{V}$.

Let's first see a lemma.

Lemma 3.3.2. We can construct an $(\alpha + 2\beta)$ -approximation [tour](#) from an (α, β) -thin tree.

Proof. Suppose we have an (α, β) -thin tree T , we want to find a *multi-subgraph* $f: \mathcal{E} \rightarrow \{0\} \cup \mathbb{N}$ such that

- (a) $f(e) \geq 1$ for all $e \in T$
- (b) $\sum_{e \in \partial^+ \{v\}} f(e) = \sum_{e \in \partial^- \{v\}} f(e)$ for all $v \in \mathcal{V}$.

We can still define an LP as follows.

$$\begin{aligned} \min \quad & \sum_{e \in \mathcal{E}} f(e)d(e) \\ & f(e) \geq 1 \quad \forall e \in T \\ & \sum_{e \in \partial^+ S} f(e) \geq |T \cap \partial^- S| \quad \forall \emptyset \neq S \subsetneq \mathcal{V} \\ & f \geq 0. \end{aligned}$$

Claim. The above LP is exact in the sense that if we have an LP solution f , we can get a [tour](#) of cost $\sum_{e \in \mathcal{E}} f(e)d(e)$.

Proof. This is just like [max-flow min-cut theorem](#). ⊛

Now, let y be

$$y_e = \begin{cases} 1 + 2\beta x_e, & \text{if } e \in T; \\ 2\beta x_e, & \text{if } e \notin T, \end{cases}$$

and the goal is to show y_e is feasible to the LP. But the only non-trivial constraints we need to check is $\sum_{e \in \partial^+ S} y_e \geq |T \cap \partial^- S|$. This follows from

$$\sum_{e \in \partial^+ S} y_e \geq 2\beta \sum_{e \in \partial^+ S} x_e = \beta \sum_{e \in \partial S} x_e \geq |T \cap \partial S| \geq |T \cap \partial^- S|.$$

Hence, y is a feasible solution of the LP, so we get a [tour](#) with cost $\sum_{e \in \mathcal{E}} y_e d(e)$, which is just

$$\sum_{e \in \mathcal{E}} y_e d(e) = \sum_{e \in T} d(e) + 2\beta \cdot \text{OPT}_{\text{LP}} \leq (\alpha + 2\beta) \text{OPT}_{\text{LP}}$$

since T is itself (α, β) -thin, which proves the result. ■

We see that [Problem 3.3.1](#) boils down to finding an (α, β) -thin tree. To do this, we'll show that by randomly sampling a [spanning tree](#), it'll be a [thin](#) tree with high probability. But the argument is non-trivial, and turns out that the number of small cuts (approximate min-cuts) is important, so we now look into this.

3.3.3 Number of Small Cuts

Given an undirected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ and a weight function $x: \mathcal{E} \rightarrow \mathbb{R}^+$, denote λ to be the minimum edge-connectivity, i.e.,

$$\lambda := \min_{\emptyset \neq S \subsetneq \mathcal{V}} \sum_{e \in \partial S} x(e),$$

we want to ask how many S achieves the value λ , i.e., how many edge min-cuts are there? It's a well-known fact that the number of the min-cuts are n^2 (or $\binom{n}{2}$ to be exact), which is tight. Now, we're interested in approximate min-cuts, or α -cuts: given $\alpha \in \mathbb{N}$, we ask that how many α -cuts S are there where an α -mincuts is defined as cuts which achieves $\sum_{e \in \partial S} x(e) \leq \alpha \cdot \lambda$? The following theorem answers this.

Theorem 3.3.2. For all $\alpha \in \mathbb{N}$, there are at most $2n^{2\alpha}$ α -mincuts.

Lecture 12: ATSP with Random Spanning Tree

Let's prove [Theorem 3.3.2](#).

10 Oct. 10:30

Proof of Theorem 3.3.2. We first see a randomized algorithm which solves the α -mincuts problem.

Algorithm 3.4: Small α -Mincuts – Karger's Algorithm [[Kar93](#)]

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, a weight function $x: \mathcal{E} \rightarrow \mathbb{R}^+$, α

Result: An α -mincut S

```

1 while  $|\mathcal{V}| > 2\alpha$  do
2    $e \leftarrow \text{rand}(\mathcal{E}, x_e)$  // Sample w.p.  $x_e$ 
3    $\text{contract}(\mathcal{G}, e)$  // new  $x$  is the sum of multi-edges'  $x$ 
4  $S \leftarrow \text{rand-subset}(\mathcal{V})$  //  $|\mathcal{V}| = 2\alpha$ 
5  $S \leftarrow \text{uncontract}(S)$ 
6 return  $S$ 
```

Now, if S is an α -mincut, we're interested in the probability of S is outputted from [Algorithm 3.4](#).

Remark. Observe that if $e = (u, v)$ is contracted where $u \in S$ while $v \notin S$, then S will definitely not be outputted.

Denote $\Pr(S \text{ survives when } |\mathcal{V}| = i) =: P_i$, then if S survived until $|\mathcal{V}| = i$, we see that S is still an α -mincut, and in the current \mathcal{G} , by considering a single vertex as a potential mincut, we have

$$\frac{\sum_{e \in \partial S} x_e}{\alpha} \leq \lambda \leq \mathbb{E}_{v \in \mathcal{V}} \left[\sum_{e \in \partial \{v\}} x(e) \right] = \frac{2}{i} \sum_{e \in \mathcal{E}} x(e) \Rightarrow P_i = 1 - \frac{\sum_{e \in \partial S} x(e)}{\sum_{e \in \mathcal{E}} x(e)} \geq 1 - \frac{2\alpha}{i} = \frac{i - 2\alpha}{i}$$

since $P_i = 1 - \Pr(S \text{ does not survive when } |\mathcal{V}| = i)$, while the latter event happens when one of the boundary edges in ∂S is picked to be contracted. We then see that

$$\begin{aligned}
\Pr(S \text{ is outputted}) &\geq P_n \cdot P_{n-1} \cdots P_{2\alpha+1} \cdot \frac{1}{2^{2\alpha}} \\
&= \left(\frac{n-2\alpha}{n} \right) \left(\frac{n-1-2\alpha}{n-1} \right) \cdots \left(\frac{1}{2\alpha+1} \right) \cdot 2^{-2\alpha} \\
&= \frac{(2\alpha)!}{n(n-1) \cdots (n-2\alpha+1)} \cdot 2^{-2\alpha} \\
&\geq \frac{1}{(2n)^{2\alpha}},
\end{aligned}$$

then the result follows because the probability should sum to 1. ■

3.3.4 Random Thin Spanning Tree

Now, we're ready to show that a randomly sampled **spanning tree** will be a **thin** tree with high probability. Recall the **spanning tree polytope**, with our **ATSP polytope** for **Problem 3.3.1**. Then given an optimal x of **Equation 3.3**, define $y_{uv} = x_{uv} + x_{vu}$ and $z = \frac{n-1}{n}y$ for **Equation 3.1**.³

Remark. z is feasible for the **spanning tree polytope**.

Proof. Since we have scaled down y , hence $\sum_{e \in \mathcal{E}} z_e = n - 1$. While for the second constraint, all $\emptyset \neq S \subsetneq \mathcal{V}$, we have

$$\sum_{e \in E(S)} x_e = \sum_{v \in S} \sum_{e \in \partial^+ \{v\}} x(e) - \sum_{e \in E(S, \bar{S})} x(e) \leq |S| - 1,$$

so z is feasible since the above sum doesn't care about direction as well, and we even decrease it a bit by down-scaling. \circledast

Now, we can sample a random **spanning tree** T' by using the **randomized pipage rounding** on z . Specifically, we have the following algorithm.

Algorithm 3.5: Thin Spanning Tree – Randomized Pipage-Rounding

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, weight $w: \mathcal{E} \rightarrow \mathbb{R}^+$, solution z of **Equation 3.1**

Result: A minimum **spanning tree** T

```

1  $T' \leftarrow \text{Randomized-Pipage-Rounding}(\mathcal{G}, w, z)$ 
2  $T \leftarrow \emptyset$ 
3 for  $e = \{u, v\} \in T'$  do                                // Using the cheaper edge between  $(u, v)$  and  $(v, u)$ 
4   if  $w((u, v)) \leq w((v, u))$  then
5      $T \leftarrow T \cup \{(u, v)\}$ 
6   else
7      $T \leftarrow T \cup \{(v, u)\}$ 
8 return  $T$ 
```

We now show that T obtained from **Algorithm 3.5** is indeed a **thin** tree.

Lemma 3.3.3. T output from **Algorithm 3.5** satisfies $\alpha = 2$ for (α, β) -thinness with probability at least $1/2$.

Proof. Since

$$\mathbb{E}[d(T)] \leq \frac{n-1}{n} \text{OPT}_{\text{LP}}$$

since we have $z_{uv} \cdot \min(d(u, v), d(v, u))$ v.s. $x_{uv}d(u, v) + x_{vu}d(v, u)$. We then see that with probability at least $1/2$ by simply using **Markov's inequality**, hence the first condition of **thin** tree is satisfied with probability at least $1/2$. \blacksquare

Lemma 3.3.4. T output from **Algorithm 3.5** satisfies $\beta = 12 \log n / \log \log n$ for (α, β) -thinness with probability at least $1 - n^{-1}$.

Proof. Given any β , we want to bound the probability that there's one $\emptyset \neq S \subsetneq \mathcal{V}$ which violates the condition. To do this, let

$$C_i := \{S \subseteq \mathcal{V} \mid z(\partial S) \in [i\lambda, (i+1)\lambda)\},$$

where we have

$$\lambda = \min_{\emptyset \neq S \subseteq \mathcal{V}} z(\partial S) = 2 \cdot \frac{n-1}{n} \geq 1.$$

Now, recall that **Algorithm 3.5** uses **Algorithm 3.2**, hence $e \in T'$ satisfies the **negative correlation**, which essentially allows us to prove **Theorem 3.2.2**. Specifically, we have

³Notice that the summation over y is exactly n , but we want the sum to be $n - 1$, hence we scale it down.

- (a) for all $e \in \mathcal{E}$, $\Pr(e \in T') := z_e$,
 (b) for all $\emptyset \neq E \subseteq \mathcal{E}$, $\Pr(E \subseteq T') \leq \prod_{e \in E} z_e$,

and with this, the [Chernoff bound-like concentration](#) states that

$$\Pr(|T' \cap S| > \beta \cdot z(\partial S)) \leq \left(\frac{e}{\beta}\right)^{\beta \cdot z(\partial S)}.$$

Then, we have

$$\begin{aligned} \Pr(\exists S: |T \cap \partial S| > \beta z(\partial S)) &\leq \sum_{i=1}^{\infty} \Pr(\exists S \in C_i: |T \cap \partial S| > \beta z(\partial S)) \\ &\leq \sum_{i=1}^{\infty} (2n)^{2(i+1)} \cdot \left(\frac{e}{\beta}\right)^{\beta \cdot i\lambda} \\ &\leq \sum_{i=1}^{\infty} (2n)^{2(i+1)} \cdot \left(\frac{e}{\beta}\right)^{\beta \cdot i}, \end{aligned}$$

where we drop λ since $\lambda \geq 1$ as we have shown, which is an even-weaker bound. From the concentration bound, we see that by taking $\beta = c \cdot \log n / \log \log n$ for some constant c ,

$$\begin{aligned} \left(\frac{e}{\beta}\right)^{\beta \cdot i} &= \left(\frac{e \log \log n}{c \log n}\right)^{\frac{c \log n}{\log \log n} \cdot i} \\ &= \exp\left(\frac{c \cdot i \cdot \log n}{\log \log n} (1 + \log \log \log n) - c - \log \log n\right) \\ &\leq \exp\left(\frac{c \cdot i \cdot \log n}{\log \log n} \left(-\frac{\log \log n}{2}\right)\right) \\ &= \exp\left(-\frac{ci}{2} \log n\right) \\ &= n^{-6i}, \end{aligned}$$

where the inequality holds when n is large, and the last equality is obtained from letting $c := 12$. Then, we see that

$$\Pr(\exists S: |T \cap \partial S| > \beta z(\partial S)) \leq \sum_{i=1}^{\infty} (2n)^{2(i+1)} \cdot \left(\frac{e}{\beta}\right)^{\beta \cdot i} \leq \sum_{i=1}^{\infty} (2n)^{2(i+1)} \cdot n^{-6i} \leq \frac{1}{n},$$

as desired. ■

Theorem 3.3.3. T output from [Algorithm 3.5](#) is a $(2, 12 \log n / \log \log n)$ -thin tree with probability at least $1/2 - n^{-1}$.

Proof. Combining [Lemma 3.3.3](#) and [Lemma 3.3.4](#) and using a union bound argument, we have the desired result. ■

In all, we have the following.

Algorithm 3.6: [Asymmetric TSP](#) – Randomized Construction

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, weight $w: \mathcal{E} \rightarrow \mathbb{R}^+$, solution z of [Equation 3.1](#)

Result: A [tour](#) T

- 1 $T' \leftarrow \text{Thin-Spanning-Tree}(\mathcal{G}, w, z)$
 - 2 $T \leftarrow \text{Thin-Tree-to-Tour}(T')$
 - 3 **return** T
-

We finally have the following.

Theorem 3.3.4. Algorithm 3.6 is an $O(\log n / \log \log n)$ -approximation algorithm with probability at least $1/2 - n^{-1}$.

Proof. By combining Theorem 3.3.3 and Lemma 3.3.2, we will obtain a $(2 + 24 \log n / \log \log n)$ -approximation tour, proving the result. ■

3.4 Symmetric Traveling Salesman Problem

Let's now look at a simpler version of Problem 3.3.1.

Problem 3.4.1 (Symmetric TSP). Given a complete graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ and a distance function $d: \mathcal{E} \rightarrow \mathbb{R}^+$ satisfying the triangle inequality. *Symmetric TSP* asks to find a tour (a_0, \dots, a_k) which minimizes $\sum_{i=0}^{k-1} d(a_{i-1}, a_i)$.

3.4.1 Christofides-Serdyuko Algorithm

We first see a simple heuristic algorithm achieves 1.5-approximation ratio of Problem 3.4.1 due to Christofides [Chr76] and Serdyukov [Ser78], discovered independently. Remarkably, this simple heuristic algorithm achieves nearly the best approximation ratio we know for more than 40 years, and the SOTA result achieves $(1.5 - 10^{-36})$ -approximation ratio [KKG21].

Algorithm 3.7: Symmetric TSP – Christofides-Serdyuko Algorithm [Chr76; Ser78]

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, a distance function $d: \mathcal{E} \rightarrow \mathbb{R}^+$

Result: A tour T

```

1  $T \leftarrow \text{MST}(\mathcal{G}, d)$                                 // Compute a minimum spanning tree
2  $O \leftarrow \{v \in T: \deg(v) \text{ is odd}\}$ 
3  $M \leftarrow \text{Min-Matching}(O, d)$                         // Compute a minimum matching
4  $T \leftarrow T \cup M$ 
5 return  $T$ 

```

Remark. We see that line 2 and line 3 solves the degree problem we have.

Proof. Explicitly, since we want a T to be (connected and) Eulerian, given a spanning tree, the only problematic part is the odd degree vertices, and hence we can just match them to solve the problem. ⊛

Clearly, Algorithm 3.7 runs in polynomial time, and we're interested in bounding the approximation ratio.

Theorem 3.4.1. Algorithm 3.7 is a 1.5-approximation algorithm.

Proof. Denote any optimal solution of Problem 3.4.1 by T^* , which is an optimal tour. Then we simply observe that $d(T') \leq d(T^*)$ and $d(M) \leq d(T^*)/2$, and we have

$$d(T) = d(T') + d(M) \leq 1.5 \cdot d(T^*),$$

proving the claim. ■

However, we get more insight by looking at the LP relaxation of Problem 3.4.1, and in fact, the recent improvement is based on looking into the corresponding LP formulation.

3.4.2 Symmetric TSP Polytope

We now analyze the approximation ratio via LP formulation of Algorithm 3.7. Indeed, since there are no directions now, we may follow the same strategy of how we solve Problem 3.3.1, i.e., we first define

the **symmetric TSP polytope** via a simple reduction from the **asymmetric TSP polytope**.

$$\begin{aligned}
 \min \quad & \sum_{e \in \mathcal{E}} x_e d(e) \\
 \sum_{e \in \partial S} x_e & \geq 2 \quad \forall \emptyset \neq S \subsetneq \mathcal{V} \\
 \sum_{e \in \partial\{v\}} x_e & = 2 \quad \forall v \in \mathcal{V} \\
 x & \geq 0
 \end{aligned} \tag{3.4}$$

Let x be this LP optimal solution, we now try to build a **tour** based on the two-step procedure as in **Algorithm 3.7**:

- (a) finding a **spanning tree**,
- (b) fix it to be in the **symmetric TSP polytope** by finding a matching of odd degrees vertices.

Analysis

To do the analysis, observe that $x \cdot \frac{n-1}{n}$ is in the **spanning tree polytope** as in the case of **Problem 3.3.1**, which implies if we can find a **minimum spanning tree** T , then $d(T) \leq \sum_{e \in \mathcal{E}} d(x)x_e$.

Note that it's not enough to just get T , we still need a matching M . Let O be the set of odd-degree vertices w.r.t. T as defined in **line 2**. Notice that we may assume $|O|$ is even, we then define the following.

Definition 3.4.1 (*O-join*). Given $O \subseteq \mathcal{V}$ and $|O|$ even, $M \subseteq \mathcal{E}$ is *O-join* if $\deg_M(v)$ is odd when $v \in O$, and $\deg_M(v)$ is even if $v \notin O$.

With this, we define the so-called **O-join LP**.

$$\begin{aligned}
 \min \quad & \sum_{e \in \mathcal{E}} y_e d(e) \\
 y(\partial S) & \geq 1 \quad \forall S \text{ s.t. } |S \cap O| \text{ odd;} \\
 y & \geq 0.
 \end{aligned} \tag{3.5}$$

Lemma 3.4.1. The **O-join LP** is exact, i.e., if y is feasible, then there exists an **O-join** M such that $d(M) \leq \sum_{e \in \mathcal{E}} d(e)y_e$.

We omit the proof here, but if we believe **Lemma 3.4.1** is true, then we have the following.

Theorem 3.4.2. **Algorithm 3.7** is a 1.5-approximation algorithm using LP relaxation analysis.

Proof. Since both **spanning tree polytope** and the **O-join LP** are valid LP relaxation of **line 1** and **line 3**, respectively, by denoting T and M obtained from solving these two LPs respectively, from the above discussion, we know

$$d(T) \leq \frac{n-1}{n} \times \text{OPT}_{\text{LP}}$$

for the **spanning tree polytope LP** and

$$d(M) \leq \frac{1}{2} \times \text{OPT}_{\text{LP}}$$

for the **O-join LP**, combining these we have

$$d(T \cup M) = d(T) + d(M) \leq 1.5 \cdot \text{OPT}_{\text{LP}}$$

for the **STSP polytope**, as desired. ■

Lecture 13: Toward Next Step: Magic Spanning Tree Distribution

3.5 Beyond the 3/2 Barrier for STSP

12 Oct. 10:30

In this section, we'll see some ideas of the recent breakthrough on [symmetric TSP](#), which breaks the 3/2-approximation barrier by a tiny absolutely constant [KKG21].

As previously seen. If we have a solution x in the [STSP polytope](#), then $(n-1)/n \times x$ is in the [spanning tree polytope](#) and $x/2$ is in the [O-join LP](#).

The advantage we'll get is that, indeed, dividing x by 2 is wasteful, or more explicitly, we want to find the set of odd degree vertices O w.r.t. T which defines the [O-join LP](#) with the cost less than $(1/2 - \delta)x$ for some absolutely constant $\delta > 0$.

Intuition. The slackness comes from the difference between all cuts and O -odd cuts.

But observe that O depends entirely on the choice of T , and hence our goal now is to sample a [spanning tree](#) T (with the corresponding $O = O^{(T)}$) such that the feasible solution $y^{(T)}$ for the corresponding [O-join LP](#) satisfies

$$(a) \mathbb{E}_T [e \in T] = (n-1)/n \cdot x_e,$$

$$(b) \mathbb{E}_T [y_e^{(T)}] \leq (1/2 - \delta)x_e.$$

If this is the case, we see that

$$\mathbb{E} [\text{cost of TSP tour}] \leq \left(\frac{3}{2} - \delta\right) \cdot \text{OPT}_{\text{LP}},$$

which breaks the 3/2 barrier for [symmetric TSP](#).

3.5.1 Strong Assumptions

Surely, we'll not look into the most general setting, instead, we'll make some strong assumptions to help us get intuitions. Specifically, we assume that there exists a tiny constant $\epsilon \in (0, 0.01)$ such that

1. $x_e \leq \epsilon$,
2. $\sum_{e \in \partial S} x(e) \geq 2 + \epsilon$ for all non-singleton cut S .⁴

The second assumption is a huge assumption, but if we have this, we see that $x/(2+\epsilon)$ satisfies all [O-join LP](#) constraints except for the singleton cut S , and we're going to fix this. To do so, we first define a useful terminology.

Notation (Even). Given a tree T , the edge $e = (u, v)$ is *even* if both $\deg_T(u)$ and $\deg_T(v)$ are even.

Then, we see that if we can sample a T such that

$$(a) \mathbb{E}_T [e \in T] = (n-1)/n \cdot x_e,$$

$$(b) \mathbb{E}_T [e \text{ is even}] \geq \delta,$$

then we'll have

$$y_e^{(T)} = \begin{cases} x_e/(2+\epsilon), & \text{if } e \text{ is even;} \\ x_e/2, & \text{otherwise,} \end{cases}$$

which is always in the [O-join polytope](#) since if e is [even](#), then the [O-join LP](#) constraint is irrelevant in this case, and hence

$$\mathbb{E} [y_e^{(T)}] \leq \delta \frac{x_e}{2+\epsilon} + (1-\delta) \frac{x_e}{2} < \left(\frac{1}{2} - \frac{\delta\epsilon}{4}\right) x_e,$$

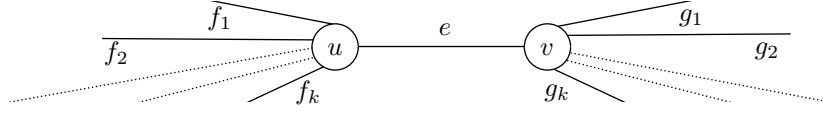
we again get the slackness we want. Now, the goal is to find such a [spanning tree](#) distribution, which is the tricky part.

⁴This means $2 \leq |S| \leq n-2$.

3.5.2 Characterization of Spanning Tree Distribution

To find such a distribution, we start by characterizing some properties which are necessary for any distribution satisfies the two conditions above. In particular, we care more about the second condition, i.e., $\mathbb{E}_T[e \text{ is even}] \geq \delta$, since the first one is quite easy to satisfy.

Fix a $(u, v) \in \mathcal{E}$, we want to make sure that $\Pr_T(\deg_T(u) \text{ and } \deg_T(v) \text{ are even}) \geq \delta$. Let f_i and g_i be the edges incident to u and v , respectively, and by abusing the notations, we also let f_i and g_i be the indicator variables indicating whether f_i appears in the tree T or not.



Note. Notice that $\deg_{\mathcal{G}}(u) = \deg_{\mathcal{G}}(v)$ since in **symmetric TSP**, \mathcal{G} is complete.

Now, define $d_u := \deg_T(u) = e + f_1 + \dots + f_k$ and $d_v := \deg_T(v) = e + g_1 + \dots + g_k$, we know that $d_u, d_v \geq 1$, and to satisfy the first condition, i.e., $\mathbb{E}_T[e \in T] = (n-1)/n \cdot x_e$, we have

$$\mathbb{E}[d_u] = \mathbb{E}[e + f_1 + \dots + f_k] = \mathbb{E}[d_v] = \mathbb{E}[e + g_1 + \dots + g_k] = \frac{n-1}{n} \cdot 2$$

since in the **symmetric TSP polytope** requires that $\sum_{e \in \partial\{v\}} x_e = 2$. We're now interested in characterizing the probability density function of d_u , which we now know it's value is at least 1 and the mean is around 2.

Log-Concavity

Let's first introduce some definition.

Definition 3.5.1 (Log-concave). Let a be an integer-valued random variable, then the distribution of a is *log-concave*^a if

$$\Pr(a = i) \geq \sqrt{\Pr(a = i-1) \Pr(a = i+1)}.$$

^aIf we take the log on both sides, we'll get a concave function.

We see that **Definition 3.5.1** can be equivalently characterize as

$$\Pr(a = i+1) \leq \Pr(a = i) \cdot \left(\frac{\Pr(a = i)}{\Pr(a = i-1)} \right).$$

Intuition (Unimodality). If a distribution is **log-concave**, going from $a = i$ to $a = i-1$, the probability decrease by a factor of α , then when considering $i+1$, it'll decrease even faster. This property is called the *unimodality*.

Example (Binomial distribution). The *binomial distribution* is **log-concave**. Furthermore, it's almost an *if and only if* condition in our case.

Proof. If we look at d_u , we can think of it as the distribution of getting heads when flipping biased coins independently with different probability each time. ⊛

With this interpretation, the distribution of d_u is like

- (a) minimum value 1,
- (b) mean 2,
- (c) essentially **binomial**,

hence it shouldn't behave too crazily. We can now prove the following.

Claim. $\Pr(d_u = 2) \geq 1/5$.

Proof. Let $\Pr(d_u = 2) =: b < 1/5$, and let $a := \Pr(d_u = 1)$, then $a \geq (1 - 1/5)/2 = 2/5$ since otherwise $\Pr(d_u \geq 3) > 2/5$ and $\mathbb{E}[d_u] > 2$, contradiction. Then, we see that the probability (ratio) gap between a and b is approximately $1/2$, from [log-concavity](#), the probability sum will just not sum up to 1, contradiction again.^a \otimes

^aIf a is large, then b need to be small, this large gap with [log-concavity](#) will make sure the same argument follows.

Also, by the similar argument, we have the following.

Claim. $\Pr(d_u \geq 2) \geq 1/2$.

Conditioning

Here, we'll see that some good properties hold after conditioning. Explicitly, let E be the event that $d_u + d_v = 4$, by the same argument as above where we now know that $d_u + d_v$ has minimum value of 2 with mean nearly 4, we have $\Pr(E) \geq 1/10$. Now, we're going to consider the probability of $\Pr(d_u = i)$ conditioning on E , i.e., $\Pr(d_u = i \mid E)$.

Remark. After conditioning, it's still [log-concave](#).

We see that if $\Pr(d_u = 2 \mid E) \geq 10\delta$, then we're done since

$$\Pr(d_u = d_v = 2) \geq \Pr(E) \cdot \Pr(d_u = 2 \mid E) \geq \frac{1}{10} \cdot 10\delta = \delta.$$

3.5.3 Toward a Contradiction

What we want is *almost* the case, since if $\Pr(d_u = 2 \mid E) \leq 10\delta$, from $10\delta < 1/3$ and the [log-concavity](#) and the only value d_u and d_v can take is 1, 2, and 3, we know that 2 can't be a mode. Hence, we see that

- 1 is the mode: $\Pr(d_u = 3 \mid E) \leq 10\delta$,
- 3 is the mode: $\Pr(d_u = 1 \mid E) \leq 10\delta$,

since we're now assuming $\Pr(d_u = 2 \mid E) \leq 10\delta$. We see that since d_u and d_v are symmetric, we may just assume

- $\Pr(d_u \leq 2 \mid E) \leq 20\delta$,
- $\Pr(d_v \geq 2 \mid E) \leq 20\delta$.

We're now going to show that this leads to a contradiction, and hence what we want is indeed the case, i.e., $\Pr(d_u = 2 \mid E) \geq 10\delta$, leading to $\Pr(d_u = d_v = 2) \geq \delta$.

Stochastic Dominance

Intuitively, stochastic dominance states that if $a + b$ is big, then a is big. From the above discussion with this intuition, we have

- $\Pr(d_u \leq 2 \mid d_u + d_v \geq 4) \leq 20\delta$,
- $\Pr(d_v \geq 2 \mid d_u + d_v \leq 3) \leq 20\delta$.

Note. From these two bound, we're almost saying something like $d_u + d_v$ is positively correlated with d_u , and is also positively correlated with d_v .

With this intuition, we see that

$$\Pr(d_u \leq 2 \text{ and } d_v \geq 2) \leq 20\delta$$

since the original two probability bounds' events are disjoint, so only one will happen. Then, we have

$$\begin{aligned}\mathbb{E}[d_u \cdot d_v] &= \mathbb{E}[d_u \cdot d_v \mid d_v = 1] \Pr(d_v = 1) + \mathbb{E}[d_u \cdot d_v \mid d_v \geq 2] \Pr(d_v \geq 2) \\ &\geq \Pr(d_v = 1) + \mathbb{E}[3 \cdot d_v \mid d_v \geq 2] \Pr(d_v \geq 2) - O(\delta) \\ &\geq 5 - O(\delta),\end{aligned}\tag{3.6}$$

where in the first inequality, in the first term, we drop $\mathbb{E}[d_u \cdot d_v \mid d_v = 1]$ naively since this is always greater than 1, and for the second term, we almost have $d_u \geq 3$ given $d_v \geq 2$ from the following intuition.

Intuition. If we first consider $\delta = 0$, then given $d_v \geq 2$, we know that from the second bound, $d_u + d_v \not\leq 3$, i.e., $d_u + d_v \geq 4$. From the first bound, this further implies $d_u \not\leq 2$, i.e., $d_u \geq 3$.

Also, for the second inequality, since $\Pr(d_v = 1) \leq 1/2$ and

$$\mathbb{E}[d_v] = \Pr(d_v = 1) + \mathbb{E}[d_v \mid d_v \geq 2] \Pr(d_v \geq 2) \approx 2,$$

the whole sum is minimized when $\Pr(d_v = 1) = 1/2$, and we get $5 - O(\delta)$.

Remark. We see that although d_u, d_v is nearly 2 as we know, but their product is at least 5.

Negative Correlation

Consider calculating $\mathbb{E}[d_u \cdot d_v]$ as follows,

$$\begin{aligned}\mathbb{E}[d_u \cdot d_v] &= \mathbb{E}\left[\left(e + \sum_i f_i\right)\left(e + \sum_i g_i\right)\right] \\ &= \mathbb{E}[e^2] + \sum_i (\mathbb{E}[e] \mathbb{E}[f_i] + \mathbb{E}[e] \mathbb{E}[g_i]) + \sum_{ij} \mathbb{E}[f_i] \mathbb{E}[g_j] \\ &= \mathbb{E}[e] + \sum_i (\mathbb{E}[e] \mathbb{E}[f_i] + \mathbb{E}[e] \mathbb{E}[g_i]) + \sum_{ij} \mathbb{E}[f_i] \mathbb{E}[g_j] \\ &\leq \epsilon + \mathbb{E}[d_u] \mathbb{E}[d_v] \\ &\leq 4 + \epsilon,\end{aligned}\tag{3.7}$$

where we use the assumption that $x_e \leq \epsilon$. We see that Equation 3.6 and Equation 3.7 gives us a contradiction.

In all, as discussed in subsection 3.5.1, we arrive at the following theorem.

Theorem 3.5.1. Under the assumptions, any spanning tree distribution satisfies log-concavity, conditioning property, stochastic dominance property, and also negative correlation property, the $3/2$ barrier of symmetric TSP can be broken.

3.5.4 Magic Spanning Tree Distribution

To see the complexity about the general cases, we now state the spanning tree distribution we'll need in the general case. There are actually three relevant distributions: given $x \in [0, 1]^{|E|}$ in the spanning tree polytope, let a distribution μ of a random spanning tree T , i.e., let $\mu(T) = \Pr(T \text{ is sampled})$.

Strongly Rayleigh Distribution

Definition. Given a distribution $\mu(T)$ and $p(z_1, \dots, z_m) := \sum_T \mu(T) \prod_{e \in T} z_e$.

Definition 3.5.2 (Real-stable). We say p is *real-stable* if $\mu(T) \in \mathbb{R}$ for all T , and for all $z_1, \dots, z_m \in \mathbb{C}$ such that $\text{Im}(z_i) > 0$ for all i , we have $p(z_1, \dots, z_m) \neq 0$.

Definition 3.5.3 (Strongly Rayleigh distribution). If μ induces a [real-stable](#) p , then μ is a *strongly Rayleigh distribution*.

Remark (Closure). If $p(z_1, \dots, z_n)$ is [real-stable](#), so are

- (a) $p(z_1, z_1, z_3, \dots, z_n)$,
- (b) $p(a, z_2, z_3, \dots, z_n)$ for all $a \in \mathbb{R}$,
- (c) $\partial p / \partial z_1$,
- (d) etc.

Turns out that there are lots of distributions are [strongly Rayleigh](#), and it's indeed a very good choice in our case since it satisfies all [log-concavity](#), [conditioning property](#), [stochastic dominance property](#), and also [negative correlation](#).

Max-Entropy Distribution

Nevertheless, among all [strongly Rayleigh distributions](#), we want to choose a most random one. This suggests we look into the notion of entropy. Denote ST be the set of all [spanning trees](#), consider the following *maximum entropy LP*:

$$\begin{array}{ll}
 \max & \sum_{T \in \text{ST}} \mu(T) \log \frac{1}{\mu(T)} \\
 & \sum_{T \in \text{ST}} \mu(T) = 1 \\
 & \sum_{T \ni e} \mu(T) = x_e \quad \forall e \in \mathcal{E} \\
 & \mu \geq 0.
 \end{array}
 \quad \Leftrightarrow \quad
 \begin{array}{ll}
 \min & \sum_{T \in \text{ST}} \mu(T) \log \mu(T) \\
 & \sum_{T \in \text{ST}} \mu(T) = 1 \\
 & \sum_{T \ni e} \mu(T) = x_e \quad \forall e \in \mathcal{E} \\
 & \mu \geq 0.
 \end{array}
 \tag{3.8}$$

Solving this induces the following.

Definition 3.5.4 (Max-entropy distribution). The optimal solution μ for the [maximum entropy LP](#) is the *max-entropy distribution*.

Observe that the [maximum entropy LP](#) has exponentially many variables, but at least this is a convex program since $x \log x$ is a convex function, and we can indeed approximately solve it in polynomial time with only polynomially many $T \in \text{ST}$ has non-zero probability, i.e., the size of the support of μ is in polynomial.

λ -Uniform Distribution

Finally, we have the following.

Definition 3.5.5 (λ -uniform distribution). A distribution $\mu(T)$ is λ -uniform if there exists a $\lambda \in (\mathbb{R}^+ \cup \{0\})^n$ such that $\mu(T) = \prod_{e \in T} \lambda_e / M$ for some M .

Note (Uniform distribution). When $\lambda_e = 1$ for all $e \in \mathcal{E}$, then we just have a uniform distribution over ST.

Now, we're going to see how these three distributions relate to each other. A technical lemma is the following.

Proposition 3.5.1. A [max-entropy distribution](#) is always [\$\lambda\$ -uniform](#) for some λ .

Proof. We can take the **Lagrangian dual** of the **maximum entropy LP**, the optimality condition (i.e., **KKT condition**) will give us the result. ■

More interestingly, we have the following.

Theorem 3.5.2 (Matrix tree theorem). A λ -uniform distribution μ for ST, then the induced p is **real-stable**, i.e., μ is **strongly Rayleigh**.

Combining **Proposition 3.5.1** and **Theorem 3.5.2**, we have the following.

Corollary 3.5.1. A **max-entropy distribution** is **strongly Rayleigh**.

We see that it's enough to find a **max-entropy distribution**, and as noted before, this can be done via solving the **maximum entropy LP** in polynomial time!

Chapter 4

Semidefinite Programming and Lasserre Hierarchy

Lecture 14: Semidefinite Programming

In this chapter, we'll talk about the Lasserre hierarchy (equivalently [Sum of Squares](#)), a good reference is Rothvoß's lecture notes [\[Rot13\]](#). 19 Oct. 10:30

4.1 Semidefinite Programming

To start with, we first introduce [semidefinite programming](#) and first develop some useful tools related to this.

4.1.1 Positive Semidefinite Matrix

In this section, an $n \times n$ matrix A is usually symmetric with real entries. In such a case, we have the following theorem.

As previously seen ([Spectrum theorem](#)). Given an $n \times n$ real and symmetric matrix A ,

- (a) There exists n real eigenvalues $\lambda_1 \geq \dots \geq \lambda_n$.
- (b) There exists n eigenvectors v_1, \dots, v_n which form an orthonormal basis.

This implies we can write $A = \sum_{i=1}^n \lambda_i v_i v_i^\top$.^a

^aConversely, if $A = \sum_{i=1}^n \lambda_i v_i v_i^\top$, then λ_i will be the eigenvalues with v_i being the corresponding eigenvector.

We now introduce the notion of [positive semidefinite](#) matrices, which is the building block of [semidefinite programmings](#).

Definition 4.1.1 (Positive semidefinite). A matrix A is *positive semidefinite* (PSD), denote as $A \succeq 0$, if for all $x \in \mathbb{R}^n$,

$$x^\top A x = \sum_{ij} x_i A_{ij} x_j \geq 0.$$

Notation. The set of real and symmetric matrices is denoted as \mathbb{S}^n , and the set of [PSD](#) matrices is denoted as \mathbb{S}_+^n .

An equivalent characterization is given by the following.

Lemma 4.1.1. $A \succeq 0$ if and only if all eigenvalues of A is non-negative.

Proof. If $\lambda_n < 0$, we have $v_n^\top A v_n = \lambda_n < 0$. On the other hand, for all $x = \sum_{i=1}^n \lambda_i v_i$, we know that $x^\top A x = \sum_i \alpha_i^2 \lambda_i \geq 0$. ■

Example. Covariance matrix, identity matrix are PSD. Also, given any V , VV^\top is PSD as well.

Proof. For VV^\top , we see that for all $x \in \mathbb{R}^n$, we have $x^\top (VV^\top) x = \langle V^\top x, V^\top x \rangle \geq 0$. ⊗

We see that there is a deeper connection between a PSD matrix and the form of VV^\top for some V , which indeed form another equivalent characterization of PSD matrices.

Lemma 4.1.2. A matrix X is PSD if and only if $X = VV^\top$ for some $V \in \mathbb{R}^{n \times k}$.

Proof. We already see that VV^\top is PSD. Now, given $X \succeq 0$, we can write $X = \sum_{i=1}^n \lambda_i v_i v_i^\top = VV^\top$ where

$$V := \text{diag} \left(\left\{ \sqrt{\lambda_1}, \dots, \sqrt{\lambda_n} \right\} \right) \begin{bmatrix} v_1 & v_2 & \dots & v_n \end{bmatrix}.$$

Remark. Given any two PSD A, B and $\alpha, \beta \geq 0$, $\alpha A + \beta B \succeq 0$.

4.1.2 Semidefinite Programming

Recall that the LP with variables $x \in \mathbb{R}^n$ is in the form of

$$\begin{aligned} \max \quad & \langle c, x \rangle \\ & \langle a_i, x \rangle \leq b_i \quad i = 1, \dots, m \\ & x \geq 0, \end{aligned}$$

with input $c, a_i \in \mathbb{R}^n$ and $b_i \in \mathbb{R}$. We can generalize LP to a much broader class of optimization problem by making vectors as matrices, which leads to the so-called **semidefinite programming**.

Definition 4.1.2 (Semidefinite programming). The *semidefinite programming (SDP)* with variables $X \in \mathbb{S}^n$ is in the form of

$$\begin{aligned} \max \quad & \langle C, X \rangle \\ & \langle A_i, X \rangle \leq b_i \quad i = 1, \dots, m \\ & X \succeq 0, \end{aligned}$$

with input $C, A_i \in \mathbb{S}^n$ and $b_i \in \mathbb{R}$.

Remark. We define the inner product between $A, B \in \mathbb{S}^n$ as

$$\langle A, B \rangle := \sum_{i,j \in [n]} A_{ij} B_{ij} = \text{tr}(AB).$$

To see that SDP is a generalization of LP, we have the following.

Lemma 4.1.3. SDP captures LP.

Proof. Given an instance of LP, i.e., given input $c, a_i \in \mathbb{R}^n$ and $b_i \in \mathbb{R}$, then consider defining $C := \text{diag}(c)$, $A_i := \text{diag}(a_i)$. Then the corresponding SDP is exactly equal to the given LP. ■

Unlike LP, where we can solve it exactly in polynomial time. Here, there's some pathological instances which cause the solution of an SDP exponential to the input size.¹ Nevertheless, we have the following.

Theorem 4.1.1. Most of SDPs can be solved in polynomial time.

¹This is the so-called bit-complexity problem.

4.1.3 Max Cut

We can imagine, [SDP](#) is usually regarded as a continuous optimization problem, and we now see one application of [SDP](#) in approximation algorithm on a combinatorial optimization problem.

Problem 4.1.1 (Max cut). Given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, find a cut $S \subseteq \mathcal{V}$ which maximizes $|\partial S|$.

Remark. There are various ways of achieving $1/2$ -approximation for [max cut](#), including greedy, local-search and also LP approaches. But we can prove that $1/2$ is tight if we only use the above method, i.e., we can't even improve this approximation ratio a bit.

Let $[n] := \mathcal{V}$ given that $n := |\mathcal{V}|$, and denote the variables being x_i for $i = 1, \dots, n$ be 1 if $i \in S$, -1 if $i \notin S$. Then the following programming captures [max cut](#) when optimizing over $x_i \in \mathbb{R}$:

$$\begin{aligned} \max \quad & \sum_{(i,j) \in \mathcal{E}} (1 - x_i x_j) / 2 \\ \text{s.t.} \quad & x_i^2 = 1 \quad \forall i \in [n]. \end{aligned}$$

Remark. This is a quadratic programming.

To solve this, we relax $x_i \in \mathbb{R}$ to $u_i \in \mathbb{R}^n$, we have

$$\begin{aligned} \max \quad & \sum_{(i,j) \in \mathcal{E}} (1 - \langle u_i, u_j \rangle) / 2 \\ \text{s.t.} \quad & \|u_i\|_2^2 = 1 \quad \forall i \in [n]. \end{aligned}$$

Now, from [Lemma 4.1.2](#), we see that $X \succeq 0 \Leftrightarrow X = VV^\top$ for some V . This suggests that the above relaxed programming is a [SDP](#) with $V = \begin{bmatrix} u_1^\top & u_2^\top & \dots & u_n^\top \end{bmatrix}$ for $u_i^\top \in \mathbb{R}^k$ such that

$$\begin{aligned} \max \quad & \sum_{(i,j) \in \mathcal{E}} (1 - X_{ij}) / 2 \quad \forall i \in [n] \\ \text{s.t.} \quad & X_{ii} = 1 \\ & X \succeq 0 \end{aligned} \tag{4.1}$$

where we let $X_{ij} = \langle u_i, u_j \rangle$. Since this is a relaxation for [max cut](#), we know that $\text{OPT}_{\text{SDP}} \geq \text{OPT}$.

Just like how we do LP relaxation, we first solve [Equation 4.1](#) to get X , and round the solution back to get ± 1 values for x_i to obtain a feasible solution for [max cut](#). A naive rounding algorithm is the following.

Algorithm 4.1: Max Cut – Randomized Rounding

Data: A connected graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, a solution X for [Equation 4.1](#)

Result: A cut $S \subseteq \mathcal{V}$

```

1  $V \leftarrow \text{Eigendecomposition}(X)$  //  $X = VV^\top$ 
2  $\begin{bmatrix} u_1^\top & u_2^\top & \dots & u_n^\top \end{bmatrix} \leftarrow V$ 
3  $g \leftarrow \text{rand}(S^{n-1})$  // Choose a random direction
4  $S \leftarrow \{i \in \mathcal{V} : \langle g, u_i \rangle \geq 0\}$  // Choose one side
5 return  $S$ 
```

We see that given $(i, j) \in \mathcal{E}$, denote the contribution to the [SDP](#) as S_{ij} , and we have

$$S_{ij} = \frac{1 - \langle u_i, u_j \rangle}{2}.$$

On the other hand, the expected contribution from (i, j) to [Algorithm 4.1](#), denoted as P_{ij} , is

$$P_{ij} := \Pr((i, j) \text{ is cut edge}) = \Pr(i, j \text{ is separated by } g).$$

Now, the only thing we need to do is to find the ratio between P_{ij} and S_{ij} . And indeed, this ratio is proved by Goemans-Williamson [[GW95](#)].

Notation (Goemans-Williamson constant). The *Goemans-Williamson constant* α_{GW} is

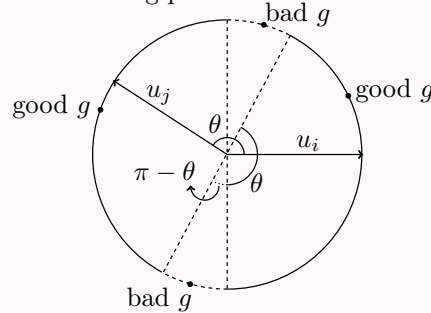
$$\frac{2}{\pi} \min_{0 \leq \theta \leq \pi} \frac{\theta}{1 - \cos \theta}.$$

Remark. $\alpha_{\text{GW}} \approx 0.878\dots$ Moreover, we can instead write α_{GW} as

$$\alpha_{\text{GW}} = \frac{2}{\pi} \min_{-1 \leq a \leq 1} \frac{\arccos(a)}{1 - a}.$$

Lemma 4.1.4. For all $(i, j) \in \mathcal{E}$, $P_{ij} \geq \alpha_{\text{GW}} \cdot S_{ij}$.

Proof. This can be seen from the following picture.



We're interested in what choice of g will not separate u_i and u_j . By drawing lines orthogonal to u_i and u_j , we divide the plane into above four regions. We see that if g lies in the dotted region, then the two inner products $\langle g, u_i \rangle$ and $\langle g, u_j \rangle$ will have the same sign, hence they're not separated.

Now, since $P_{ij} = 2\theta/2\pi = \theta/\pi$, and recall that $S_{ij} = (1 - \langle u_i, u_j \rangle)/2$, so the worst configuration producing the smallest ratio between P_{ij} and S_{ij} is

$$\frac{P_{ij}}{S_{ij}} \geq \min_{a \in [-1, 1]} \frac{\arccos(a)/\pi}{(1 - a)/2} = \alpha_{\text{GW}},$$

where $a := \langle u_i, u_j \rangle$. This proves the result. ■

Theorem 4.1.2 ([GW95]). [Algorithm 4.1](#) is an α_{GW} -approximation algorithm in expectation.

Proof. Given S outputted from [Algorithm 4.1](#), we see that

$$\frac{\mathbb{E}[|S|]}{\text{OPT}_{\text{SDP}}} = \frac{\sum_{(i,j) \in \mathcal{E}} P_{ij}}{\sum_{(i,j) \in \mathcal{E}} S_{ij}} \geq \alpha_{\text{GW}}$$

from [Lemma 4.1.4](#), which proves the result. ■

Lecture 15: Lasserre Hierarchy and Max Cut

4.2 Lasserre Hierarchy

24 Oct. 10:30

The Lasserre hierarchy is a systematic procedure to strengthen a relaxation for an optimization problem by adding additional variables and [SDP](#) constraints. In the last years this hierarchy moved into the focus of researchers in approximation algorithms as they obtain relaxations have provably nice properties.

4.2.1 Local Distributions

Firstly, recall **max cut** and its IP formulation and the **SDP** relaxation. We're interested in whether there's something *between* the IP and this **SDP** relaxation?

$$\begin{array}{ccc} \max & \frac{1}{4} \sum_{(i,j) \in \mathcal{E}} \|u_i - u_j\|_2^2 & \rightarrow \dots \rightarrow \max & \frac{1}{4} \sum_{(i,j) \in \mathcal{E}} (x_i - x_j)^2 \\ \text{(IP)} & \|u_i\|_2^2 = 1 \quad \forall i \in [n], & \text{(SDP)} & x_i \in \{\pm 1\} \quad \forall i \in [n]. \end{array}$$

The upshot is that the **SDP** solutions are kind of telling us the second moment information. In order to see this, instead of optimizing over ± 1 , we now optimize over $\{0, 1\}$.

$$\begin{array}{ccc} \max & \sum_{(i,j) \in \mathcal{E}} \|u_i - u_j\|_2^2 & \rightarrow \dots \rightarrow \max & \sum_{(i,j) \in \mathcal{E}} (x_i - x_j)^2 \\ \text{(IP)} & \|u_\emptyset\|_2^2 = 1 \quad \forall i \in [n] & \text{(SDP)} & x_i \in \{0, 1\} \quad \forall i \in [n]. \\ & \langle u_i, u_\emptyset \rangle = \|u_i\|_2^2 \quad \forall i \in [n] & & \end{array}$$

Now, we ask the following question.

Problem. If $u_\emptyset, \{u_i\}_{i \in [n]}$ are feasible to $\{0, 1\}$ -SDP, what does it tell us about the integral solutions?

Answer. If $\{u_i\}_{i \in [n]}, u_\emptyset$ is 1-dimensional, then this solution $\{u_i\}_{i \in [n]}, u_\emptyset$ encodes a cut $S \subseteq \mathcal{V}$ such that

$$S = \{i \in [n] \mid u_i = 1\},$$

in which case we can get

$$\langle u_i, u_j \rangle = \mathbb{1}(i, j \in S)$$

for all $i, j \in [n]$. Now, if $\{u_i\}_{i \in [n]}, u_\emptyset$ is **not** 1-dimensional, then we hope that we can view the solution $\{u_i\}_{i \in [n]}, u_\emptyset$ is encoding a *distribution* \mathcal{D} over cuts $S \subseteq \mathcal{V}$, in which case, we can think of

$$\langle u_i, u_j \rangle = \mathbb{E}_{S \sim \mathcal{D}} [\mathbb{1}(i, j \in S)]$$

for all $i, j \in [n]$, i.e., a covariance matrix. But sadly, this is not true in this exact form since a **PSD matrix** doesn't always stand for a covariance matrix of some distribution over $\{0, 1\}$ -valued assignments. ⊛

To get at least some versions of what we want, we first introduce a special kind of distribution called **2-local distribution**.

Definition 4.2.1 (2-local distribution). A *2-local distribution* is a set of distributions consisting of

- \tilde{P}_i : distribution over $\{0, 1\}$ -assignments for $X_i = \mathbb{1}(i \in S)$ for all $i \in [n]$;
- \tilde{P}_{ij} : distribution over $\{0, 1\}$ -assignments for (X_i, X_j) for all $(i, j) \in [n] \times [n]$,

which satisfies the **2-local consistency**.

Definition 4.2.2 (2-local consistency). The set of distributions \tilde{P}_i and \tilde{P}_{ij} is *2-local consistent* if for all $i, j \in [n]$ and for all $\theta \in \{0, 1\}$,

$$\tilde{P}_i(X_i = \theta) = \sum_{\theta' \in \{0, 1\}} \tilde{P}_{ij}(X_i = \theta, X_j = \theta') = \tilde{P}_{ij}(X_i = \theta).$$

Example. If $\tilde{P}_i(X_i = \theta) = 1$ and $\tilde{P}_{ij}(X_i = \theta) = 0$, then this set is not a **2-local distribution**.

Now, consider the $\{0, 1\}$ -SDP with local probabilities. The **2-local** variables are $\{\tilde{P}_i\}_{i \in [n]} \cup \{\tilde{P}_{ij}\}_{i, j \in [n]}$

with $\{v_i\}_{i \in [n]}$ and v_\emptyset . Then the SDP for **max cut** is defined as

$$\begin{aligned} \max \quad & \sum_{(i,j) \in \mathcal{E}} \|u_i - u_j\|_2^2 \\ & \langle u_i, u_\emptyset \rangle = \tilde{P}_i(X_i = 1) \quad \forall i \in [n] \\ & \langle u_i, u_j \rangle = \tilde{P}_{ij}(X_i = X_j = 1) \quad \forall i, j \in [n] \end{aligned}$$

Remark. Technically, we should also introduce another distribution $\tilde{P}_\emptyset(X_\emptyset = 1)$, and \tilde{P}_{ij} is defined for all $(i, j) \in ([n] \cup \{\emptyset\}) \times ([n] \cup \{\emptyset\})$. In this case, the **SDP** constraint reduces to

$$\langle u_i, u_j \rangle = \tilde{P}_{ij}(X_i = X_j = 1)$$

for all $(i, j) \in ([n] \cup \{\emptyset\}) \times ([n] \cup \{\emptyset\})$.

We first investigate the objective. We see that

$$\begin{aligned} \|u_i - u_j\|_2^2 &= \|u_i\|_2^2 + \|u_j\|_2^2 - 2\langle u_i, u_j \rangle \\ &= \tilde{P}_i(X_i = 1) + \tilde{P}_j(X_j = 1) - 2\tilde{P}_{ij}(X_i = X_j = 1) \\ &= \tilde{P}_{ij}(X_i = 1) + \tilde{P}_{ij}(X_j = 1) - 2\tilde{P}_{ij}(X_i = X_j = 1) = \tilde{P}_{ij}(X_i \neq X_j). \end{aligned}$$

Also, observe that

$$\langle u_i, u_j \rangle = \mathbb{E}_{\tilde{P}_{ij}}[x_i x_j],$$

hence we create a matrix $M \in \mathbb{R}^{([n] \cup \{\emptyset\}) \times ([n] \cup \{\emptyset\})}$ such that $M_{ij} = \mathbb{E}_{\tilde{P}_{ij}}[x_i x_j]$, i.e., $M = UU^\top$. In this case, the original constraint implies that M is **PSD**, hence overall, the **SDP** becomes

$$\begin{aligned} \max \quad & \sum_{(i,j) \in \mathcal{E}} \tilde{P}_{ij}(X_i \neq X_j) \\ & \{\tilde{P}_i\} \cup \{\tilde{P}_{ij}\} \text{ is a } \mathbf{2\text{-local distribution}} \\ (\mathcal{P}) \quad & M = \left(\mathbb{E}_{\tilde{P}_{ij}}[X_i X_j] \right)_{i,j \in [n] \cup \{\emptyset\}} \succeq 0, \end{aligned}$$

where we call this **SDP** \mathcal{P} .

We see that the notion of **2-local distribution** can be generalized to arbitrary number R , i.e., we can now look at the so-called **R -local distribution**.

Definition 4.2.3 (Local distribution). The set of distributions $\{\tilde{P}_A\}_{A \subseteq [n] \cup \{\emptyset\}, |A| \leq R}$ is an **R -local distribution** if for all $A, B \subseteq [n] \cup \{\emptyset\}$ with $|A \cup B| \leq R$, for all $C \subseteq A \cup B$ and $\theta_i \in \{0, 1\}$,

$$\tilde{P}_C(X_i = \theta_i \forall i \in C) = \tilde{P}_A(X_i = \theta_i \forall i \in C) = \tilde{P}_B(X_i = \theta_i \forall i \in C).$$

Note. Notice that we can also define the generalized version of **2-local consistency**, but we just encoded this in **Definition 4.2.3**.

Now, the **R -local version** of \mathcal{P} becomes

$$\begin{aligned} \max \quad & \sum_{(i,j) \in \mathcal{E}} \tilde{P}_{ij}(X_i \neq X_j) \\ & \{\tilde{P}_A\}_{|A| \leq R} \text{ is an } \mathbf{R\text{-local distribution}} \\ \text{Lass}_R(\mathcal{P}) \quad & M = \left(\mathbb{E}_{\tilde{P}_{A \cup B}} \left[\prod_{i \in A \cup B} X_i \right] \right)_{A,B} \succeq 0, \end{aligned}$$

where we call this **SDP** $\text{Lass}_R(\mathcal{P})$, which is how we define Lasserre hierarchy.

Note. Notice that $M \in \mathbb{R}^{\binom{[n]}{\leq R/2} \times \binom{[n]}{\leq R/2}}$

Lecture 16: Lasserre Hierarchy Continued

After defining $\text{Lass}_R(\mathcal{P})$, we now analyze what kind of properties this hierarchy has.

26 Oct. 10:30

Remark. Up to this time, we have seen the following.

- (a) $\text{Lass}_R(\mathcal{P})$ is a convex program with $2^R n^{O(R)}$ variables and constraints.
- (b) We can solve this in $n^{O(R)}$ time.
- (c) $\text{Lass}_2(\mathcal{P})$ is equivalent to a basic SDP, and $\text{Lass}_n(\mathcal{P})$ is equivalent to an IP.

4.2.2 Probabilistic Consequences

Consider

- $\text{Var}_{\tilde{P}_i} [X_i] = \mathbb{E}_{\tilde{P}_i} [X_i^2] - \mathbb{E}_{\tilde{P}_i} [X_i]^2$
- $\text{Cov}_{\tilde{P}_{ij}} [X_i, X_j] = \mathbb{E}_{\tilde{P}_{ij}} [X_i X_j] - \mathbb{E}_{\tilde{P}_i} [X_i] \mathbb{E}_{\tilde{P}_j} [X_j]$.

We see that we can do a conditioning: let $\tilde{P} := \{\tilde{P}_A\}_{|A| \leq R}$ be an R -local distribution. Now, fix $S \subseteq [n]$, $|S| = t$, let $\alpha_S = \{0, 1\}^{|S|}$, condition on \tilde{P} , we get

$$\tilde{P}' = \tilde{P} \mid X_S \leftarrow \alpha_S,$$

where $X_S \leftarrow \alpha_S$ means $X_i \leftarrow \alpha_i$ for all $i \in S$.

Remark. \tilde{P}' is a $(R - t)$ -local distribution.

Proof. For all $A \subseteq [n]$, $|A| \leq R - t$, we have

$$\tilde{P}'_A(X_A = \theta_A) = \frac{\tilde{P}(X_A = \theta_A, X_S = \alpha_S)}{\tilde{P}(X_S = \alpha_S)}.$$

⊛

Apart from this, we also see that

- (a) \tilde{P}' is $(R - t)$ -wise locally consistent.
- (b) If \tilde{P} was $\text{Lass}_R(\mathcal{P})$, \tilde{P}' is feasible for $\text{Lass}_{R-t}(\mathcal{P})$.

Lemma 4.2.1 (Conditioning reduces variance). For all $i, j \in [n]$,

$$\text{Var}_{\tilde{P}_i} [X_i] - \mathbb{E}_{X_j \sim \tilde{P}_j} [\text{Var}_{\tilde{P}_{ij}} [X_i \mid X_j]] \geq 4 \text{Cov}_{\tilde{P}_{ij}} [X_i, X_j]^2.$$

Proof. From of law of total variance, we have

$$\text{Var}_{\tilde{P}_i} [X_i] - \mathbb{E}_{X_j} [\text{Var}_{\tilde{P}} [X_i \mid X_j]] = \text{Var}_{\tilde{P}_j} [\mathbb{E}_{\tilde{P}_j} [X_i \mid X_j]].$$

Now, let $P_i = \tilde{P}_i(X_i = 1)$, $P_j = \tilde{P}_j(X_j = 1)$, $P_{ij} = \tilde{P}_{ij}(X_i = 1, X_j = 1)$, we have

$$\begin{aligned}
\text{Var}_{X_j} [\mathbb{E}_{\tilde{P}} [X_i | X_j]] &= \mathbb{E}_{X_j} [\mathbb{E}_{\tilde{P}} [X_i | X_j]^2] - (\mathbb{E}_{X_j} [\mathbb{E}_{\tilde{P}} [X_i | X_j]])^2 \\
&= \tilde{P}_j(X_j = 1) \cdot \frac{\mathbb{E}_{\tilde{P}} [X_i X_j]^2}{\tilde{P}_j(X_j = 1)^2} + \tilde{P}_j(X_j = 0) \cdot \frac{\mathbb{E}_{\tilde{P}} [X_i(1 - X_j)]^2}{\tilde{P}_j(X_j = 0)^2} - \mathbb{E}_{\tilde{P}} [X_i]^2 \\
&= \frac{P_{ij}^2}{P_j} + \frac{(P_i - P_{ij})^2}{1 - P_j} - P_i^2 \\
&= \frac{1}{P_j(1 - P_j)} (P_{ij}^2(1 - P_j) + (P_i - P_{ij})^2 \cdot P_j - P_i^2 P_j(1 - P_j)) \\
&= \frac{(P_{ij} - P_i P_j)^2}{P_j(1 - P_j)} \\
&= \frac{(\mathbb{E}_{\tilde{P}} [X_i X_j] - \mathbb{E}_{\tilde{P}_i} [X_i] \mathbb{E}_{\tilde{P}_j} [X_j])^2}{\mathbb{E} [X_j^2] - \mathbb{E} [X_j]^2} \\
&= \frac{\text{Cov}_{\tilde{P}} [X_i, X_j]^2}{\text{Var}_{\tilde{P}_j} [X_j]}.
\end{aligned}$$

Since X_i are 0-1 variable, the variance in the denominator is less than 1/4, hence we finally have

$$\text{Var}_{\tilde{P}_i} [X_i] - \mathbb{E}_{X_j} [\text{Var}_{\tilde{P}} [X_i | X_j]] = \text{Var}_{X_j} [\mathbb{E}_{\tilde{P}} [X_i | X_j]] \geq 4 \cdot \text{Cov}_{\tilde{P}} [X_i, X_j]^2.$$

Corollary 4.2.1. Suppose $\tilde{P} = \{\tilde{P}_A\}_{|A| \leq R}$ is an R -local distribution which is $\text{Lass}_R(\mathcal{P})$ feasible, then

$$\mathbb{E}_{j \sim [n]} \mathbb{E}_{X_j \sim \tilde{P}_j} [\mathbb{E}_{i \sim [n]} [\text{Var}_{\tilde{P}_i} [X_i]] - \mathbb{E}_{i \sim [n]} [\text{Var}_{\tilde{P}_{ij}} [X_i | X_j]]] \geq 4 \mathbb{E}_{i, j \sim [n]} [\text{Cov}_{\tilde{P}_{ij}} [X_i, X_j]^2].$$

Furthermore, given $a \in \mathbb{R}^+$, either one of the following will happen.

- (a) $\mathbb{E}_{i, j \sim [n]} [\text{Cov}_{\tilde{P}_{ij}} [X_i, X_j]^2] \leq a.$
- (b) $\exists j \in [n], \theta_j \in \{0, 1\}, \tilde{P}' := \tilde{P} \mid X_j \leftarrow \theta_j$ satisfies $\mathbb{E}_{i \sim [n]} [\text{Var}_{\tilde{P}_i} [X_i]] - \mathbb{E}_{i \sim [n]} [\text{Var}_{\tilde{P}'} [X_i]] \geq 4a.$

Proof. We first prove the first statement. Lemma 4.2.1 gives a point-wise inequality, taking the expectation on both sides with the **dominant convergence theorem**, we have

$$\mathbb{E}_{i, j \sim [n]} [\text{Var} [X_i] - \mathbb{E}_{X_j} [\text{Var} [X_i | X_j]]] \geq 4 \mathbb{E}_{i, j \sim [n]} [\text{Cov} [X_i, X_j]^2],$$

with the fact that

$$\mathbb{E}_{i, j \sim [n]} [\text{Var} [X_i] - \mathbb{E}_{X_j} [\text{Var} [X_i | X_j]]] = \mathbb{E}_{j \sim [n]} \mathbb{E}_{X_j} [\mathbb{E}_{i \sim [n]} [\text{Var} [X_i]] - \mathbb{E}_{i \sim [n]} [\text{Var} [X_i | X_j]]],$$

hence conclude the first part. A probabilistic argument proves the *either-or* statement. ■

Remark. Corollary 4.2.1 says that either we have a small covariance, or we can reduce it by a lot.

Theorem 4.2.1. Suppose $\tilde{P} = \{\tilde{P}_A\}_{|A| \leq R}$ is an R -local distribution which is $\text{Lass}_R(\mathcal{P})$ feasible and $R \geq 1/\epsilon^4 + 2$, then there exists $S \subseteq [n]$ such that $|S| \leq 1/\epsilon^4$, $\alpha_S \in \{0, 1\}^{|S|}$, and $\tilde{P}' := \tilde{P} \mid X_S \leftarrow \alpha_S$, we have

$$\mathbb{E}_{i, j \sim [n]} [\text{Cov}_{\tilde{P}'} [X_i, X_j]^2] \leq \frac{\epsilon^4}{4}.$$

Moreover, S and α_S can be found in $\text{poly}(n, 1/\epsilon)$.

Proof. We actually have a constructive proof, i.e., we directly give an algorithm which runs in $\text{poly}(n, 1/\epsilon)$ and find the desired i, j .

Algorithm 4.2: Theorem 4.2.1 – Construction

Data: $\tilde{P}, \epsilon > 0$
Result: \tilde{P}' with expected covariance smaller than $\epsilon^4/4, S$

```

1  $\ell \leftarrow 0, \tilde{P}^{(\ell)} \leftarrow \tilde{P}, S \leftarrow \emptyset$ 
2 for  $\ell = 0, 1, \dots, 1/\epsilon^4$  do
3   if  $\mathbb{E}_{i,j} [\text{Cov}_{\tilde{P}^{(\ell)}} [X_i, X_j]^2] \leq \epsilon^4/4$  then
4     return  $\tilde{P}^{(\ell)}$  //  $\tilde{P}^{(\ell)} = \tilde{P} \mid X_S \leftarrow \alpha_S, S$ 
5   else
6     Find  $j_{\ell+1} \in [n] \setminus S, \theta_{\ell+1} \in \{0, 1\}$  // Guaranteed in Corollary 4.2.1
7      $\tilde{P}^{(\ell+1)} \leftarrow \tilde{P}^{(\ell)} \mid X_{j_{\ell+1}} \leftarrow \theta_{\ell+1}$ 
8      $S \leftarrow S \cup \{j_{\ell+1}\}$ 
```

To analyze Algorithm 4.2, observe that if Algorithm 4.2 returns, then we have a desired property, so we only need to ensure it'll meet the condition in line 3 in $1/\epsilon^4$ iterations. Now, for a **local distribution** Q , let $\text{Var}[Q] := \mathbb{E}_{i \sim [n]} [\text{Var}_Q [X_i]]$ and $\text{Cov}[Q] := \mathbb{E}_{i,j \sim [n]} [\text{Cov}_Q [X_i, X_j]]$. We see that we only fail if in every iteration, we reach line 5, i.e., $\text{Cov}[\tilde{P}^{(\ell)}] \leq \epsilon^4/4$ for all ℓ . But from Corollary 4.2.1, we know that the $\tilde{P}^{(\ell+1)}$ we find will have the property that

$$\text{Var}[\tilde{P}^{(\ell-1)}] - \text{Var}[\tilde{P}^{(\ell)}] \geq 4 \cdot \epsilon^4/4 = \epsilon^4 \Rightarrow \text{Var}[\tilde{P}^{(\ell)}] \leq \text{Var}[\tilde{P}^{(\ell-1)}] - \epsilon^4.$$

By telescoping, we have

$$\text{Var}[\tilde{P}^{(1/\epsilon^4)}] \leq \text{Var}[\tilde{P}^{(0)}] - \frac{1}{\epsilon^4} \cdot \epsilon^4 = \text{Var}[\tilde{P}] - 1 \leq \frac{1}{4} - 1 < 0,$$

a contradiction, and hence we must terminate, finishing the proof. ■

Remark. Theorem 4.2.1 says that suppose we have a **local distribution** over n variables with sufficient large locality. Then turns out that there's a small subset of variables, if we fix them, they'll almost determine all other variables.

Finally, we have the following algorithm.

Algorithm 4.3: Max Cut – PTAS

Data: A dense graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ with $|\mathcal{E}| \geq \epsilon n^2, \epsilon > 0$
Result: A cut S

```

1  $R \leftarrow 1/\epsilon^4 + 2$ 
2  $\tilde{P} := \{\tilde{P}_A\}_{|A| \leq R} \leftarrow \text{Solve}(\text{Lass}_R(\mathcal{P}))$ 
3  $\tilde{P}' \leftarrow \text{Reduce-Variance}(\tilde{P})$  //  $\mathbb{E}_{i,j \sim [n]} [\text{Cov}_{\tilde{P}'} [X_i, X_j]^2] \leq \epsilon^4/4$ 
4
5 // Rounding
6 for  $i \in \mathcal{V}$  do
7    $\lambda_i \leftarrow \text{Ber}(\tilde{P}'(X_i = 1))$ 
8  $S \leftarrow \{i \in \mathcal{V} : \lambda_i = 1\}$ 
9 return  $S$ 
```

Remark. The rounding method in Algorithm 4.3 (i.e., line 4) is ridiculously simple compare to Algorithm 4.1! \tilde{P}' basically tells you everything.

We indeed have the following guarantee.

Theorem 4.2.2 (PTAS for max cut). For any $\epsilon > 0$, given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ such that $|\mathcal{E}| \geq \epsilon n^2$, there exists a $(1 - 4\epsilon)$ -approximation algorithm runs in $n^{O(1/\epsilon^4)}$ -time.

We see that as long as the graph is dense enough, we can spend more and more time to get a better approximation, which is the whole point of Lasserre hierarchy.

Lecture 17: Graph Coloring

Let's first prove [Theorem 4.2.2](#).

31 Oct. 10:30

Proof of Theorem 4.2.2. The running time for [Algorithm 4.3](#) is clear. Denote $p_i = \tilde{P}'(X_i = 1)$, then we see that the expected fraction of edges cuts is

$$\begin{aligned} & \mathbb{E}_{(i,j) \in \mathcal{E}} [\text{Pr}_{\text{ALG}}(X_i \neq X_j)] \\ &= \mathbb{E}_{X_S} \mathbb{E}_{(i,j) \in \mathcal{E}} [p_i + p_j - 2p_i p_j] \\ &\geq \mathbb{E}_{X_S} \mathbb{E}_{(i,j) \in \mathcal{E}} [p_i + p_j - 2\mathbb{E}_{\tilde{P}'}[X_i X_j] - 2|p_i p_j - \mathbb{E}_{\tilde{P}'}[X_i X_j]|] \\ &= \underbrace{\mathbb{E}_{(i,j) \in \mathcal{E}} [\mathbb{E}_{\tilde{P}}[X_i] + \mathbb{E}_{\tilde{P}}[X_j] - 2\mathbb{E}_{\tilde{P}}[X_i X_j]]}_{\text{Lass}} - \underbrace{2\mathbb{E}_{X_S} \mathbb{E}_{(i,j) \in \mathcal{E}} [p_i p_j - \mathbb{E}_{\tilde{P}'}[X_i X_j]]}_{\text{Err}}. \end{aligned}$$

Recall that previously, we only have control on $\text{Cov}_{\tilde{P}'}^2 = \text{Cov}^2(\tilde{P} \mid X_S \leftarrow \alpha_S)$, which is over the whole $i, j \sim [n]$. But now the error term (the second term) is only over $(i, j) \sim \mathcal{E}$, hence we define

$$\text{Cov}_{\mathcal{E}}^2 [\tilde{P} \mid X_S \leftarrow \alpha_S] = \mathbb{E}_{X_S} \mathbb{E}_{(i,j) \sim \mathcal{E}} [\text{Cov}[X_i, X_j \mid S_S]^2].$$

Claim. $\text{Cov}_{\mathcal{E}}^2[\tilde{P} \mid S] \leq \text{Cov}[\tilde{P} \mid X_S \leftarrow \alpha_S]/\epsilon \leq \epsilon^3$.

Proof. We see that

$$\begin{aligned} n^2 \cdot \text{Cov}^2 [\tilde{P} \mid X_S \leftarrow \alpha_S] &= \text{Cov}_{\Sigma}^2 [\tilde{P} \mid X_S \leftarrow \alpha_S] \\ &\geq \text{Cov}_{\mathcal{E}, \Sigma}^2 [\tilde{P} \mid X_S \leftarrow \alpha_S] = m \cdot \text{Cov}_{\mathcal{E}}^2 [\tilde{P} \mid X_S \leftarrow \alpha_S], \end{aligned}$$

where subscript Σ is when we replace the expectation by summation in the covariance. With the fact that $m \leftarrow \epsilon n^2$, we're done. \otimes

Now, since we know that $\text{Lass} \geq \text{OPT} \geq 1/2$, we have

$$\mathbb{E}_{(i,j) \in \mathcal{E}} [\text{Pr}_{\text{ALG}}(X_i \neq X_j)] \geq \text{Lass} - 2\epsilon \geq \text{Lass}(1 - 4\epsilon),$$

finishing the proof. \blacksquare

4.3 Graph Coloring

Return to the [SDP](#), first, we introduce a new definition.

Definition 4.3.1 (Coloring). Given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, a (valid) coloring $\chi: \mathcal{V} \rightarrow [c]$ is a function χ such that for all $(i, j) \in \mathcal{E}$, $\chi(i) \neq \chi(j)$.

Now, consider the following problem.

Problem 4.3.1 (Graph coloring). Given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, find a coloring $\chi: \mathcal{V} \rightarrow [c]$ while minimizing c .

Before trying to solve the [graph coloring](#), we note that it's trivial to get n -coloring (by using different color for every node). But in fact, this is the best we can do: [graph coloring](#) is extremely hard!

Theorem 4.3.1. For all $\epsilon > 0$, it's NP to get $n^{1-\epsilon}$ -approximation.

People start to consider some promise version of [graph coloring](#), i.e., if we directly assume that \mathcal{G} admits a c -coloring, what can we say?

- $c = 1$: \mathcal{G} has no edges, hence trivial.
- $c = 2$: \mathcal{G} is bipartite, so we can color the graph alternatively, so this can be solved exactly.
- $c = 3$: We can do $\tilde{O}(n^{1/4})$ -approximation (quite shameful...)

Remark (SOTA for $c = 3$). For $c = 3$, someone showed that we can do $\tilde{O}(n^{0.199\dots})$, i.e., around $\tilde{O}^{1/5}$. Also, $\omega(1)$ -approximation is NP!

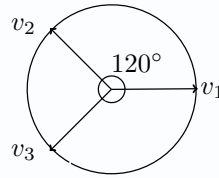
Analogous to [max cut](#), we design the following SDP relaxation of [graph coloring](#) with variables being vectors v_i for $i \in \mathcal{V}$.

$$\begin{aligned} \min \quad & 0 \\ & \langle v_i, v_i \rangle = 1 \quad \forall i \in \mathcal{V} \\ & \langle v_i, v_j \rangle = -1/2 \quad \forall (i, j) \in \mathcal{E}. \end{aligned} \tag{4.2}$$

Note. We don't have an actual objective function!

Claim. If \mathcal{G} is 3-colorable, there exists $\{v_i\}_{i \in \mathcal{V}}$ that are feasible for [Equation 4.2](#).

Proof. Consider 3 colors C_1, C_2 , and C_3 , then if i has C_j , we let i gets vector v_j with all vectors $v_j, v_{j'}$ are 120° away.



⊛

4.3.1 Independent Sets

Turns out that the [independent set](#) is highly related to solving just [Equation 4.2](#), as we'll soon see.

Definition 4.3.2 (Independent set). Given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, a set $S \subseteq \mathcal{V}$ is *independent* if for all $(i, j) \in \mathcal{E}$, either $i \notin S$ or $j \notin S$.

The notion of [independent set](#) is useful since we can transform the [graph coloring](#) problem into finding [independent sets](#) as suggests by [Lemma 4.3.1](#).

Lemma 4.3.1. For a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, then \mathcal{G} is c -colorable if and only if \mathcal{V} can be partitioned to V_1, \dots, V_c such that V_i is [independent](#).

With [Lemma 4.3.1](#), if we can find large [independent sets](#) and partition the graph into not too many of those, we're done. Note that the size of [independent sets](#) is related to the maximum degree.

Notation. We denote the maximum degree of a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ by $\Delta := \Delta(\mathcal{G}) := \max_{v \in \mathcal{V}} \deg(v)$.

Remark (Usefulness of Δ). Given Δ , we will have a trivial Δ -coloring; also, we know that we can find independent set with size $n/(\Delta + 1) = \Omega(n\Delta^{-1})$.

Proof. The coloring part is clear. As for finding independent set, we see that by randomly include one vertex to our independent set, we at most Δ vertices will be ruled out: they can't be in the independent set now. \otimes

Now, after solving Equation 4.2, notice that we only have feasibility, with the fact that the solution are not guaranteed to be perfectly aligned in exactly three vectors, hence it's a bit confusing what to do next. However, recall that it's also good enough to find a large independent set, and recall the max cut problem, where we want to maximize the number of edges crossing a cut set, which is similar to what we're trying to do here. So, inspired by which, we can round the solution, but observe the following.

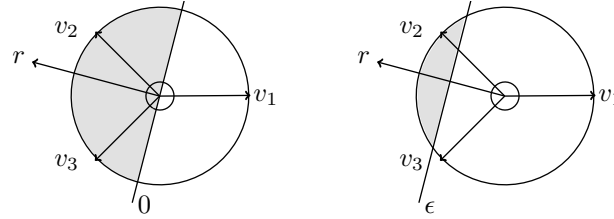


Figure 4.1: If we do the rounding as in max cut, we may end-up including more than we want, so we set up some threshold.

This suggests that we round it with threshold, i.e., consider the following algorithm.

Algorithm 4.4: Graph Coloring – Independent Set Rounding of 3-Colorable Graph

Data: A 3-colorable graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$

Result: An independent set S

- 1 $\{v_i \in \mathbb{R}^d\}_{i=1}^n \leftarrow \text{Solve}(\text{SDP})$
 - 2 $r \leftarrow \mathcal{N}(0, I_d)$ // $r_i \sim \mathcal{N}(0, 1)$
 - 3 $S(\epsilon) \leftarrow \{i \in \mathcal{V} : \langle r, v_i \rangle \geq \epsilon\}$ // $\epsilon = \sqrt{2/3 \cdot \ln \Delta}$
 - 4 $S'(\epsilon) = \{i \in S(\epsilon) : \nexists j \in S(\epsilon) \text{ s.t. } (i, j) \in \mathcal{E}\}$ // Make S independent
 - 5 **return** $S'(\epsilon)$
-

Remark. Algorithm 4.4 is the rounding algorithm of Equation 4.2 in the sense of feasibility,^a and notice that it gives us an independent set, rather than a coloring.

^aRecall that there's no objective in Equation 4.2

Lecture 18: Graph Coloring via Independent Sets Decomposition

We're now interested in how large the independent set Algorithm 4.4 outputs. To start analyzing, since the r sampled in line 2 is Gaussian, recall the following. 2 Nov. 10:30

As previously seen (Gaussian distribution). The probability density function for Gaussian distribution is

$$p(x) = \frac{1}{\sqrt{2\pi}} e^{-x^2/2},$$

and the cumulated density function is

$$\Phi(x) = \Pr(g \leq x) = \int_{-\infty}^x p(s) ds, \quad \bar{\Phi}(x) = \Pr(g \geq x) = 1 - \Phi(x).$$

From the spherical symmetry, we have $\langle r, v_i \rangle \sim \mathcal{N}(0, 1)$ for all $i \in \mathcal{V}$. Moreover, since the probability of i being in $S(\epsilon)$ is exactly $\bar{\Phi}(\epsilon)$, from the linearity of expectation, we have $\mathbb{E}[|S(\epsilon)|] = n \cdot \bar{\Phi}(\epsilon)$.

Lemma 4.3.2. $\Pr(i \notin S'(\epsilon) \mid i \in S(\epsilon)) \leq \Delta \bar{\Phi}(\sqrt{3}\epsilon)$.

Proof. Fix any $(i, j) \in \mathcal{E}$, it's sufficient to show $\Pr(j \in S(\epsilon) \mid i \in S(\epsilon)) \leq \bar{\Phi}(\sqrt{3}\epsilon)$. And from the fact that all v_j are 120° apart, we hence can write

$$v_j = -\frac{1}{2}v_i + \frac{\sqrt{3}}{2}u$$

where $\|u\| = 1$ and $u \perp v_i$. If $j \in S(\epsilon)$, then

$$\langle v_j, r \rangle \geq \epsilon \Rightarrow \underbrace{\left\langle -\frac{1}{2}v_i, r \right\rangle}_{\leq -\epsilon/2} + \left\langle \frac{\sqrt{3}}{2}u, r \right\rangle \geq \epsilon \Rightarrow \left\langle \frac{\sqrt{3}}{2}u, r \right\rangle \geq \frac{3\epsilon}{2} \Rightarrow \langle u, r \rangle \geq \sqrt{3}\epsilon.$$

Since if $u \perp v \in \mathbb{R}^d$, $\langle u, r \rangle$ and $\langle v, r \rangle$ are independent, so $\Pr(j \in S(\epsilon) \mid i \in S(\epsilon)) \leq \bar{\Phi}(\sqrt{3}\epsilon)$ as desired. ■

We can now prove that the independent set found by Algorithm 4.4 is large.

Theorem 4.3.2. Algorithm 4.4 finds an independent set of size $\Omega(n \cdot \Delta^{-1/3} \log^{-1/2} \Delta)$ for any 3-colorable \mathcal{G} .

Proof. We see that

$$\mathbb{E}[|S'(\epsilon)|] = \sum_{i \in \mathcal{V}} \underbrace{\Pr(i \in S(\epsilon)) \cdot \Pr(i \in S'(\epsilon) \mid i \in S(\epsilon))}_{\bar{\Phi}(\epsilon)} \geq \sum_{i \in \mathcal{V}} \bar{\Phi}(\epsilon) \cdot (1 - \Delta \bar{\Phi}(\sqrt{3}\epsilon)),$$

from Lemma 4.3.2. Now, observe the following.

Claim. If $x \geq 10$, $p(x)/2x \leq \bar{\Phi}(x) \leq p(x)/x$.

Proof. Since we know that

$$\frac{x}{1+x^2} \cdot p(x) \leq \bar{\Phi}(x) \leq \frac{1}{x} \cdot p(x)$$

for all x , if $x \geq 10$, $x/(1+x^2) \geq 1/2x$, hence we're done. ⊗

With the above claim, we have $\bar{\Phi}(\epsilon) \geq p(\epsilon)/2\epsilon$ with $\bar{\Phi}(\sqrt{3}\epsilon) \leq p(\sqrt{3}\epsilon)/3\epsilon$, hence

$$p(\sqrt{3}\epsilon) = \frac{1}{\sqrt{2\pi}} e^{-3 \cdot (2/3 \cdot \ln \Delta)/2} = \frac{1}{\sqrt{2\pi}} \cdot \frac{1}{\Delta} \text{ and } p(\epsilon) = \frac{1}{\sqrt{2\pi}} e^{-2/3 \cdot \ln \Delta/2} = \frac{1}{\sqrt{2\pi}} \frac{1}{\Delta^{1/3}},$$

leading to

$$\mathbb{E}[|S'(\epsilon)|] \geq \sum_{i \in \mathcal{V}} \underbrace{\left(\frac{1}{2\epsilon} \cdot p(\epsilon) \right)}_{\geq \Omega(\frac{1}{\sqrt{\ln \Delta}} \frac{1}{\Delta^{1/3}})} \cdot \underbrace{\left(1 - \Delta \cdot \frac{1}{3\epsilon} \cdot p(\sqrt{3}\epsilon) \right)}_{\geq 1/2} \geq \Omega(n \cdot \Delta^{-1/3} \ln^{-1/2} \Delta).$$

■

4.3.2 Independent Sets Decomposition

From Lemma 4.3.1, we can iteratively find large independent sets by Algorithm 4.4 as guaranteed by Theorem 4.3.2. But before we see the final algorithm, we introduce a cute trick.

Remark (Wigderson's trick). We can always 3-color $\{v\} \cup \mathcal{N}(v)$ for all $v \in \mathcal{V}$ if \mathcal{G} is 3-colorable.

Proof. Since for a 3-colorable graph, for all $v \in \mathcal{V}$, $\mathcal{G}[N(v)]$ is bipartite (all $u \in N(v)$ already links with v , so the degree will be at most 2 in $\mathcal{G}[N(v)]$). And as mentioned before, we can always 2-color a bipartite graph. And we just use another new color for v to do the 3-coloring. \otimes

Now, we see the final algorithm.

Algorithm 4.5: Graph Coloring – Independent Set Decomposition of 3-Colorable Graph

Data: A 3-colorable graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$

Result: A colored \mathcal{G}

```

1  $n_0 \leftarrow |\mathcal{V}|$ 
2
3 // Phase 1
4 while  $\Delta(\mathcal{G}) \geq n_0^{3/4}$  do
5    $v \leftarrow \arg \max_{i \in \mathcal{V}} \deg(i)$ 
6   3 colors  $\{v\} \cup N(v)$  // Wigderson's trick
7    $\mathcal{G} \leftarrow \mathcal{G}[\mathcal{V} \setminus (\{v\} \cup N(v))]$ 
8
9 // Phase 2
10 while  $\Delta(\mathcal{G}) \geq 100$  do //  $\Delta(\mathcal{G}) < n_0^{3/4}$ 
11    $S \leftarrow \text{Independent-Set}(\mathcal{G})$  //  $|S| \geq cn\Delta^{-1/3} \ln^{-1/2} \Delta$  from Theorem 4.3.2
12   1 colors  $S$ 
13    $\mathcal{G} \leftarrow \mathcal{G}[\mathcal{V} \setminus S]$ 
14
15 // Phase 3
16  $\Delta(\mathcal{G}) + 1$  colors  $\mathcal{G}$  //  $\Delta(\mathcal{G}) < 100$ 
17 return  $\mathcal{G}$ 
```

Notice that in Algorithm 4.5, whenever we do a coloring, we use a brand-new color to avoid any collision.

Theorem 4.3.3. Algorithm 4.5 is an $\tilde{O}(n^{1/4})$ -approximation algorithm for graph coloring.

Proof. We see that

- **Phase 1:** color at least $n_0^{3/4}$ vertices with 3 colors in each iteration, hence need at most $n_0^{1/4} (= 3 \cdot n_0/n_0^{3/4})$ colors.
- **Phase 2:** from Theorem 4.3.2, $|S| \geq nc n_0^{-1/4} \log^{-1/2} n_0 =: n\gamma$, then the induced graph will have vertices less than $n(1 - \gamma)$.^a Hence, we can run this at most k iteration since $n \geq 1$, i.e.,

$$1 \leq n \leq n_0(1 - \gamma)^k \leq n_0 e^{-\gamma k} \Rightarrow k \leq \frac{1}{\gamma} \ln n_0 = \frac{1}{c} n_0^{1/4} \log^{1/2} n_0 \cdot \ln n_0$$

- **Phase 3:** Clean-up phase, only uses constant amount (< 100) more colors.

In all, Algorithm 4.5 uses at most

$$3n_0^{1/4} + \frac{1}{c} n_0^{1/4} \log^{3/2} n_0 + 100 = \tilde{O}(n^{1/4})$$

colors. ■

^aNotice the different between n and n_0 : n is updating, while n_0 is the original graph size.

Remark. We can use the similar approach for small constant c , e.g., $c = 4$, $c = 5$, etc.

Chapter 5

Hardness of Approximation

Lecture 19: Complexity theory for Approximation Algorithm

Recall how we define the [combinatorial optimization](#).

7 Nov. 10:30

As previously seen. Given a set of all possible inputs \mathcal{I} for a [combinatorial optimization](#) problem P , the goal is to find $x \in X_I$ to maximize/minimize $f_I(x)$ where $f_I: X_I \rightarrow \mathbb{R}^+$ is an objective function, and X_I is a set of feasible solutions.

Now, we're going to discuss the complexity of doing approximation problem. But since the classical complexity theory is under the context of decision problems, we now try to generalize it.

5.1 Approximation Complexity

In this section, we're going to consider maximization problems primarily. Since we now care about decision problem, so given a maximization [problem](#) P with goal being finding an objective in \mathbb{R} , we have the following decision version.

Definition 5.1.1 (Decision- P). Given a [maximization problem](#) P , the *decision- P* is the decision version of P , where given an input $I \in \mathcal{I}$, $c \in \mathbb{R}^+$, finds an algorithm which output True if $\text{OPT}_I \geq c$, False otherwise.

And we have the following characterization of P and [decision- \$P\$](#) in terms of complexity class.

Definition 5.1.2. A [maximization problem](#) P is NP if [decision- \$P\$](#) is NP.

Apparently, this is not enough since what we care is the approximation version, and hence we have the following generalization.

Definition 5.1.3 (α -Gap). Given a [maximization problem](#) P with $\alpha < 1$, the α -Gap P is the [decision version](#) of α -approximating P , where given an input $I \in \mathcal{I}$ and $c \in \mathbb{R}^+$, finds an algorithm which outputs True if $\text{OPT}_I \geq c$, False if $\text{OPT}_I < \alpha c$, and anything else (don't care) otherwise.

Since we see that we may ignore some outputs, we divide the output into two different sets.

Notation. Let \mathcal{I} be a set of all inputs, Y be a set of all True inputs, N be a set of all False inputs.^a Then given an input $I \in \mathcal{I}$, output True if $I \in Y$, False if $I \in N$, anything otherwise.

^aWe have $Y \cap N = \emptyset$, while not necessarily have $N \cup Y = \mathcal{I}$.

Remark. If there is an α -approximation algorithm for P , then there is an algorithm for [\$\alpha\$ -Gap \$P\$](#) .

Proof. Given I and c , we run the α -approximation algorithm for P to get a solution with value c' .

Notice that we necessarily have

$$\alpha \text{OPT}_I \leq c' \leq \text{OPT}_I,$$

hence we can design a new algorithm which outputs **True** if $c' \geq c \cdot \alpha$, **False** otherwise. This is a correct algorithm for α -Gap P since

- If $\text{OPT}_I \geq c$, then $c' \geq \alpha \text{OPT}_I \geq \alpha c$, which is the **True** case.
- If $\text{OPT}_I < \alpha c$, then $c' \leq \text{OPT}_I < \alpha c$, which is the **False** case.

Conversely, if there is no polynomial time algorithm for α -Gap P , there is no α -approximation algorithm for P . *

Again, we have the following characterization of P and α -Gap P in terms of complexity class.

Definition 5.1.4. An α -approximating problem P is NP if the α -Gap P is NP.

5.1.1 Approximation Reduction

Finally, we briefly review the **reduction**.

As previously seen (Reduction). Given two problems P_1, P_2 , a *reduction* from P_1 to P_2 is a polynomial time algorithm R such that given an input $I_1 \in \mathcal{L}_1$, output $I_2 \in \mathcal{L}_2$ satisfying both **completeness** and **soundness**.

As previously seen (Completeness). A **reduction** from P_1 to P_2 satisfies *completeness* if it transforms an accepted input for P_1 to an accepted input for P_2 , i.e., if $I_1 \in Y_1$, then $I_2 \in Y_2$.

As previously seen (Soundness). A **reduction** from P_1 to P_2 satisfies *completeness* if it transforms a rejected input for P_1 to a rejected input for P_2 , i.e., if $I_1 \in N_1$, then $I_2 \in N_2$.

The reason why we care about **reduction** is that given a reduction R from P_1 to P_2 , if there exists an algorithm for P_2 , then we have an algorithm for P_1 . by the following.

Algorithm 5.1: Reduction

Data: Algorithm \mathcal{A}_2 for P_2 , **reduction** R from P_1 to P_2 , input $I_1 \in \mathcal{L}_1$ for P_1

Result: Decision of I_1

```

1  $I_2 \leftarrow R(I_1)$ 
2 return  $\mathcal{A}_2(I_2)$ 

```

Similarly, since we care about approximation algorithm, we can define the approximation **reduction** from α -Gap P_1 to β -Gap P_2 .

Definition 5.1.5 (Reduction). Given two problems P_1, P_2 , a *reduction* from P_1 to P_2 is a polynomial time algorithm R such that given an input $I_1 \in \mathcal{L}_1$ and c_1 , output $I_2 \in \mathcal{L}_2$ and c_2 satisfying both **completeness** and **soundness**.

Definition 5.1.6 (Completeness). A **reduction** from P_1 to P_2 satisfies *completeness* if it transforms an accepted input for P_1 to an accepted input for P_2 , i.e., if $\text{OPT}_{I_1} \leq c_1$, then $\text{OPT}_{I_2} \leq c_2$.

Definition 5.1.7 (Soundness). A **reduction** from P_1 to P_2 satisfies *completeness* if it transforms a rejected input for P_1 to a rejected input for P_2 , i.e., if $\text{OPT}_{I_1} > \alpha c_1$, then $\text{OPT}_{I_2} > \beta c_2$.

And again, given a **reduction** R if there is a polynomial time algorithm for β -Gap P_2 , then we have a polynomial time algorithm for α -Gap P_1 ; on the other hand, if there is no polynomial time algorithm for α -Gap P_1 , then there is no polynomial time algorithm for β -Gap P_2 , so there is no β -approximation algorithm for P_2 .

5.2 Probabilistically Checkable Proofs

5.2.1 Constraint Satisfaction Problem

We first study one of the most important problems in theoretical computer science, the **CSP** problem. This is important since it's the **reduction** for many important problems, and from the discussion, if we have a good algorithm for **CSP**, we automatically get lots of other problems solved.

Problem 5.2.1 (CSP). Given an input $(x_1, \dots, x_n) = X$, C_1, \dots, C_m where $C_i = (a_i, b_{i_1}, \dots, b_{i_k})$ be the set of clauses where $a_i \in \ell$, $b_{i_j} \in [n]$, the *constraint satisfaction problem* of Σ, Φ^a is to find $\sigma: X \rightarrow \Sigma$ maximizing the number of satisfied clauses, i.e., $\sigma_{a_i}(x_{b_{i_1}}, \dots, x_{b_{i_k}}) = 1$.

^a Σ is the alphabet set and $\Phi = \{\phi_1, \dots, \phi_\ell\}$ is a family of constraints where $\phi_i: \Sigma^k \rightarrow \{0, 1\}$.

There's an important distinction between problem description and problem instance. That is, the **CSP** with respect to Σ, Φ is the problem description of a class of problems, and after given some variables X and clauses C_i , it becomes a problem instance, which can be solved.

Notation (Problem description). The problem description of **CSP** with respect to Σ and Φ is denoted as $\text{CSP}(\Sigma, \Phi)$.

Notice that we can equivalently maximize the fraction instead of maximize the number of satisfied clauses, i.e., the objective is now $\# \text{satisfied clauses} / m$. It's because it's convenient to normalize the objective to be in $[0, 1]$.

Note. Notice that to represent $\phi_i: \Sigma^k \rightarrow \{0, 1\}$, it's often more convenient just to denote it as $\phi_i^{-1}(\{1\})$, i.e., the set of accepted string in Σ^k w.r.t. ϕ_i .

Example (Max-cut as CSP). **Max cut** is equivalent to $\text{CSP}(\Sigma, \Phi)$ where $\Sigma = \{0, 1\}$, $\Phi = \{\phi_1\}$ with $\phi_1 = \{01, 10\}$.

Proof. If we model **max cut** in this way, given an instance of **max cut**, i.e., given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ with n nodes, $C_i = (1, u, v)$ for $(u, v) \in \mathcal{E}$. The first entry is 1 since there are only one constraint to check whether a node is in the cut or not, and we create C_i for every edge (u, v) . *

Example (Max-2SAT as CSP). **MAX-2SAT** is equivalent to $\text{CSP}(\Sigma, \Phi)$ where $\Sigma = \{0, 1\}$, $\Phi = \{\phi_1, \dots, \phi_4\}$ with

$$\begin{aligned} \phi_1 &= \{01, 10, 11\} \Leftrightarrow (x_i \vee x_j), & \phi_2 &= \{01, 10, 00\} \Leftrightarrow (\bar{x}_i \vee \bar{x}_j), \\ \phi_3 &= \{01, 00, 11\} \Leftrightarrow (\bar{x}_i \vee x_j), & \phi_4 &= \{00, 10, 11\} \Leftrightarrow (x_i \vee \bar{x}_j). \end{aligned}$$

5.2.2 The Probabilistic Checkable Proofs Theorem

As mentioned, there's lots of **reduction** can be done between fundamental problems considered in TCS to **CSP**, including one of the most important results in hardness, the **PCP theorem**. In order to do this, we need a more fine-grained version of **Definition 5.1.3**.

Definition 5.2.1 ((c, s)-Gap). Given a maximization **problem** P with $0 < s \leq c \leq 1$, the (c, s) -Gap P is the **decision version** of α -approximating P , where given an input $I \in \mathcal{I}$ and $c \in \mathbb{R}^+$, finds an algorithm which outputs True if $\text{OPT}_I \geq c$, False if $\text{OPT}_I < s$, and anything else (don't care) otherwise.

Note. We implicitly assume that (c, s) -Gap P is only defined for P being a **CSP**, or can be **reduced** to **CSP**.

Remark. We see that by setting $s = \alpha \cdot c$, we recover [Definition 5.1.3](#) from [Definition 5.2.1](#).

Then, we have the following.

Theorem 5.2.1 (Cook-Levin theorem [[Coo71](#)]). The $(1, 1)$ -Gap 3SAT is NP-hard.

Theorem 5.2.2 (Karp [[Kar72](#)]). For all fixed $\epsilon > 0$, $(1 - \epsilon, 1 - \epsilon)$ -Gap max cut is NP-hard.

Note. The $(1, 1)$ -Gap max cut is P.

Proof. Recall that if we transform max cut into CSP, the optimal value is always 1,^a i.e., every edge is cut edge, so in this case the graph must be bipartite. This can be easily check. ⊛

^ai.e., we're not comparing to the optimal value of one instance of \mathcal{G} .

Theorem 5.2.3 (PCP theorem [[Fei+91](#); [Aro+98](#)]). There exists an $\epsilon > 0$ such that $(1, 1 - \epsilon)$ -Gap 3SAT is NP-hard.

To understand , we need to understand the class PCP. First, recall the definition of NP.

As previously seen (NP). A language $L \subseteq \{0, 1\}^*$ is in NP if there exists a Turing machine V runs in $\text{poly}(|x|)$ such that given x ,

- $x \in L$, then $\exists y$ such that $V(x, y) = 1$;
- $x \notin L$, then $\forall y$ such that $V(x, y) = 0$.

Definition 5.2.2 (PCP). The class *probabilistically checkable proofs*, or $\text{PCP}_{c,s}(r(n), q(n))$,^a is defined as $L \in \text{PCP}_{c,s}(r, q)$ if there exists a poly-time randomized Turing machine V which can only flip r coins^b and given an input x , V can look at x on q position Q_1, \dots, Q_q by $\phi_R: \{0, 1\}^q \rightarrow [0, 1]$ where

- $x \in L$, then $\exists y$ such that $\Pr_R(\phi(y_{Q_1}, \dots, y_{Q_q}) = 1) \geq c$;
- $x \notin L$, then $\forall y$ such that $\Pr_R(\phi(y_{Q_1}, \dots, y_{Q_q}) = 1) < s$.

^aWe implicitly assume that r and q depends on the length of the input $|x| = n$.

^bIt only accepts random string R with length r , i.e., is $R \in \{0, 1\}^r$.

In [Definition 5.2.2](#), the randomized Turing machine V decides both the position (Q_1, \dots, Q_q) we're allowed to access, and also a function ϕ_R which only looks at x_{Q_1}, \dots, x_{Q_q} , acting as a decider for V .

Note. Everything is decided before looking at any input.

Just like [Cook-Levin theorem](#) is the mother of all *exact hardness*, [PCP theorem](#) is the mother of all *hardness of approximation*.

Lecture 20: FGLSS Graph

With [Definition 5.2.2](#), [PCP theorem](#) is equivalent to saying the following.

9 Nov. 10:30

Theorem 5.2.4 (Equivalent [Theorem 5.2.3](#)). The [PCP theorem](#) is equivalent as saying that there exists $\epsilon > 0$ such that

$$\text{NP} = \text{PCP}_{1,1-\epsilon}(O(\log n), O(1)).$$

Proof. It's easy to see that $\text{NP} \supseteq \text{PCP}_{1,1-\epsilon}(O(\log n), O(1))$ just by considering iterating through all the possible R . Another direction worth a whole class, so we're not going to dive into that. ■

Nevertheless, if we accept that $\text{NP} = \text{PCP}_{1,1-\epsilon}(O(\log n), O(1))$, we can actually show the equivalence between [Theorem 5.2.4](#) and the [PCP theorem](#) by showing that [Theorem 5.2.4](#) implies hardness of approximation, specifically, the $(1, 1-\epsilon)$ -Gap 3SAT problem.

Firstly, from [Cook-Levin theorem](#), 3SAT is $\text{NP} = \text{PCP}_{1,1-\epsilon}(O(\log n), O(1))$ from assumption. But instead of demonstrate the [reduction](#) to 3SAT, we consider [max cut](#) instead.

Remark. Generally, between two [Gap CSPs](#) with (c_1, s_1) and (c_2, s_2) , the hardness is preserve, so we may consider [max cut](#) instead since it can be modeled as [CSP](#), and use the machinery to show the hardness for 3SAT.^a

^aFor more detailed explanation, see [Piazza](#).

Assume $q = 2$, $\psi = \{01, 10\}$, and $r = O(\log n)$. Then there exists V such that given a 3-CNF formula ϕ , it runs in $\text{poly}(|\phi|)$ and only flips r random coins $R \in \{0, 1\}^r$, which decides Q_1^R, Q_2^R such that

- if ϕ is satisfiable, $\exists y$ such that $\Pr_R(\psi(y_{Q_1^R}, y_{Q_2^R}) = 1) \geq c$;
- if ϕ is not satisfiable, $\forall y$ such that $\Pr_R(\psi(y_{Q_1^R}, y_{Q_2^R}) = 1) \leq s$.

Notice that the above event $\psi(y_{Q_1^R}, y_{Q_2^R})$ is exactly $y_{Q_1^R} \neq y_{Q_2^R}$. We see that there are at most $2^r \leq n^{O(1)}$ possible R 's, and for each R , we access exactly 2 positions, so V will access at most $N := 2 \cdot 2^r$ positions. Now, without loss of generality, we may assume that $\max_R(Q_1^R, Q_2^R) \leq N$.¹

Consider the optimization problem that finds $y \in \{0, 1\}^N$ to maximize the probability of accepting. In this viewpoint, this is just like [max cut](#) on $\mathcal{G} = ([N], \{(Q_1^R, Q_2^R) : R \in \{0, 1\}^r\})$. Namely, we find a [reduction](#) from 3SAT to (c, s) -Gap max cut.

5.3 FGLSS Graph

To see how we utilize [PCP theorem](#), we first see one example.

Problem 5.3.1 (Vertex cover). Given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, find the smallest $C \subseteq \mathcal{V}$ that covers all \mathcal{E} .

Problem 5.3.2 (Independent set). Given a graph $\mathcal{G} = (\mathcal{V}, \mathcal{E})$, find the largest $I \subseteq \mathcal{V}$ that contains no edge.

[Problem 5.3.1](#) and [Problem 5.3.2](#) are often considered together due to the following relation.

Claim. For all \mathcal{G} , $\text{OPT}_{\text{VC}}(\mathcal{G}) = |\mathcal{V}| - \text{OPT}_{\text{IS}}(\mathcal{G})$.

Proof. Observe that for all $C \subseteq \mathcal{V}$, C is a [vertex cover](#) if and only if $\mathcal{V} \setminus C$ is an [independent set](#). \otimes

5.3.1 Hardness of Vertex Cover and Independent Set

The hardness of [vertex cover](#) and [independent set](#) can be shown by using the FGLSS graph [\[Fei+96\]](#), which allows us to do [reduction](#) from $(1, s)$ -Gap 3SAT with $s < 1$, which is NP-hard from the [PCP theorem](#).

Consider the input of the $(1, s)$ -Gap 3SAT being a 3CNF formula ϕ , n variables $X = \{x_1, \dots, x_n\}$ and $2n$ literals $L = \{x_1, \bar{x}_1, x_2, \bar{x}_2, \dots\}$ with m clauses $\{C_1, \dots, C_m\}$ with three literals in each, i.e., $C_i = (\ell_{i_1} \vee \ell_{i_2} \vee \ell_{i_3})$ with $\ell_{i_j} \in L$.

Then, the goal is to find a [reduction](#) from this input to an input (in both cases, it's a graph) of α -Gap vertex cover and β -Gap independent set for some α, β . Toward this goal, we consider the so-called FGLSS graph [\[Fei+96\]](#) $\mathcal{G} = (\mathcal{V}, \mathcal{E})$ such that

- $\mathcal{V} = [m] \times \left(\{T, F\}^3 \setminus (F, F, F) \right)$ with $|\mathcal{V}| = 7m$;
- $((i, \ell_{i_1}, \ell_{i_2}, \ell_{i_3}), (j, \ell_{j_1}, \ell_{j_2}, \ell_{j_3})) \in \mathcal{E}$ if they *contradict*.

¹Since V is going to access at most N positions anyway, we can just rearrange it.

The interpretation is that each vertex (i, t_1, t_2, t_3) indicates value of $(\ell_{i_1}, \ell_{i_2}, \ell_{i_3})$, i.e., it's a partial assignment for only 3 variables in C_i .

Notation (Contradiction). If the partial assignment given by two vertices in a FGLSS graph is not consistent, we say they are *contradicting*.

Example. Given $C_1 = (\bar{x}_1 \vee x_2 \vee \bar{x}_3)$ and $C_2 = (x_3 \vee x_4 \vee x_5)$ with two vertices $v = (1, T, T, T)$ and $u = (2, T, F, F)$, they are contradicting to each other.

Proof. Since v states that \bar{x}_3 is T (x_3 is F); while u states that x_3 is T , they contradict. \circledast

This actually finishes the **reduction**, and the only thing left to do is to determine what α and β is. To do this, observe that following.

Remark. Denote $V_i := \{i\} \times (\{T, F\}^3 \setminus (F, F, F))$, we see that V_i is a clique with size 7 since they all contradict to each other.

This means that for **independent set**, $c = m$; and for **vertex cover**, $c = |\mathcal{V}| - m = 7m - m = 6m$. We first show the **completeness**.

Claim. If $\text{OPT}_{3\text{SAT}}(\phi) = 1$, then $\text{OPT}_{\text{IS}}(\mathcal{G}) \geq m$ and $\text{OPT}_{\text{VC}}(\mathcal{G}) \leq 6m$.

Proof. Since ϕ is satisfiable, then there exists $\sigma: X \rightarrow \{T, F\}$ that satisfies every C_i . Then from each V_i , choose a vertex *consistent* with σ .^a

And since they come from the same assignment σ , there are no contradiction hence no edges between these vertices, i.e., they form a **independent set**. Hence, $\text{OPT}_{\text{IS}}(\mathcal{G}) \geq m$, and $\text{OPT}_{\text{VC}}(\mathcal{G}) \leq 6m$. \circledast

^aThere are exactly one for each i .

We now show the **soundness**. In particular, we will always deal with contrapositive in this course, i.e., instead of find a bad input from a bad input, we find a good input from a good input, but backwards.

Claim. If $\text{OPT}_{3\text{SAT}}(\phi) < s$, then $\text{OPT}_{\text{IS}}(\mathcal{G}) < sm$ and $\text{OPT}_{\text{VC}}(\mathcal{G}) > (7 - s)m$.

Proof. Consider the contrapositive, i.e., we show that $\text{OPT}_{\text{IS}}(\mathcal{G}) \geq sm$ (hence $\text{OPT}_{\text{VC}}(\mathcal{G}) \leq (7 - s)m$), then $\text{OPT}_{3\text{SAT}}(\phi) \geq s$.

Let $I \subseteq \mathcal{V}$ be an **independent set** such that $|I| \geq sm$, and let $\sigma: X \rightarrow \{T, F\}$ such that for all C_i with $|I \cap V_i| = 1$,^a assign variables in C_i according to $I \cap V_i$.^b Finally, we extend it arbitrarily for unassigned variables if needed. We see that for all C_i such that $|I \cap V_i| = 1$, this assignment σ satisfies C_i , hence σ satisfies exactly $|I| \geq sm$ clauses, i.e., the normalized optimal solution for **3SAT** is $\geq sm/m = s$ as required. \circledast

^aIt can only be the case that I doesn't include vertices from some V_i , but if it does, no more than 1 can be included since V_i is a clique.

^bThis is well-defined since there are no contradictions with I being an **independent set**.

With the above discussion, we see that the **Gap** is $\beta = s < 1$, $\alpha = (7 - s)/6 > 1$. Hence, there exists a **reduction** from **(1, s)-Gap SAT** to **(7 - s)/6-Gap vertex cover** and **s-Gap independent set**.

Remark. Actually, it's also easy to check that there exists a **reduction** from **(c, s)-Gap P** to **(f - s)/(f - c)-Gap vertex cover** and **s/c-Gap independent set** for any **CSP** P with f being the number of satisfying assignments.

From this, we have the following.

Theorem 5.3.1. For all $\epsilon > 0$, there exists a **CSP** P such that **(1, ϵ)-Gap P** is NP-hard.

Corollary 5.3.1. For all $c > 0$, there exists no c -approximation algorithm for [independent set](#).

The state-of-the-art in-approximation result for [independent set](#) result is the following.

Theorem 5.3.2. For all $\epsilon > 0$, there exists no $1/n^{1-\epsilon}$ -approximation algorithm for [independent set](#).

5.4 Label Cover

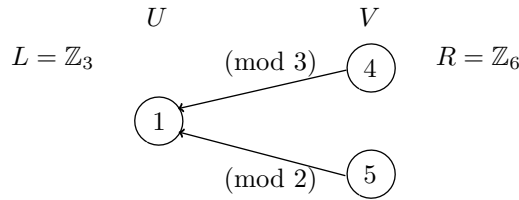
Although [PCP theorem](#) is powerful as we just saw, but there is also another useful problem to study when doing [reduction](#), the [label cover](#).

Problem 5.4.1 (Label cover). Given a bipartite graph $\mathcal{G} = (U \sqcup V, \mathcal{E})$ with label sets L (for U) and R (for V) with $|R| \geq |L|$ such that for all $e = (u, v) \in \mathcal{E}$, we have a projection $\Pi_e: [R] \rightarrow [L]$. The *label cover* problem asks for an assignment $\sigma: U \sqcup V \rightarrow L \cup R$ such that

$$\sigma|_U: U \rightarrow L, \quad \sigma|_V: V \rightarrow R,$$

which satisfies the maximum number of edge.^a

^aThe edge $e = (u, v)$ is satisfied by σ if $\Pi_e(\sigma(v)) = \sigma(u)$.



For [label cover](#), the parameters are $|U|, |V|, |\mathcal{E}|, L, R$, and we sometimes for simplicity, use L and R to also denote the size of L and R .

Remark (Baseline). There is a trivial $1/L$ -approximation algorithm.

Proof. Consider a random assignment σ such that

- for all $v \in V$, $\sigma(v)$ randomly from $[R]$;
- for all $u \in U$, $\sigma(u)$ randomly from $[L]$.

Fix $e = (u, v)$, we see that $\Pr(\Pi_e(\sigma(v)) = \sigma(u)) = 1/L$. ⊛

Theorem 5.4.1. For all $\epsilon > 0$, there exists L, R such that the $(1, \epsilon)$ -Gap [label cover](#) for L, R is NP-hard.

Proof. This is based on the [PCP theorem](#) with [parallel repetition theorem](#). ■

5.4.1 Hardness of Max k -Coverage

Recall the [max \$k\$ -coverage](#) problem.

Problem 5.4.2 (Max k -coverage). Given a [set system](#) Ω, \mathcal{S} and k , finds $\mathcal{S}' \subseteq \mathcal{S}$ such that $|\mathcal{S}'| = k$ which maximizes $|\bigcup_{S \in \mathcal{S}'} S|$.

We're going to see the hardness of the [max \$k\$ -coverage](#) problem, and our goal is to prove the following.

Theorem 5.4.2. For all $\epsilon > 0$, there is no $(3/4 + \epsilon)$ -approximation algorithm for [max \$k\$ -coverage](#).

Interestingly, the state-of-the-art result is the following.

Theorem 5.4.3. For all $\epsilon > 0$, there is no $(1 - 1/e + \epsilon)$ -approximation algorithm for [max \$k\$ -coverage](#).

By proving [Theorem 5.4.2](#), we almost prove [Theorem 5.4.3](#)! Let's first see the reduction. Given a [label cover](#) instance $\mathcal{G} = (U \sqcup V, \mathcal{E})$, L, R and $\{\Pi_e\}_{e \in \mathcal{E}}$, consider

$$\Omega := \mathcal{E} \times \{0, 1\}^L$$

such that $(e, x_1, \dots, x_L) \in \Omega$ with $|\Omega| = |\mathcal{E}| \cdot 2^L$. Then, the reduction is given by

- For all $u \in U$, $i \in [L]$, $S_{u,i} = \{(e, x_1, \dots, x_L) : e \ni u, x_i = 0\}$.
- For all $v \in V$, $i \in [R]$, $S_{v,i} = \{(e, x_1, \dots, x_L) : e \ni v, x_{\Pi_e(i)} = 0\}$.

Appendix

Appendix A

Review

A.1 Boolean Satisfaction Problem

Here, we give a quick review toward the [MAX-3SAT](#) problem.

Definition A.1.1 (Conjunctive normal form). A *conjunctive normal form* (CNF) formula is a conjunction φ of one or more boolean clauses on x_1, x_2, \dots, x_n with boolean valued $\{0, 1\}$. Explicitly, φ is in CNF if

$$\varphi(x_1, x_2, \dots, x_n) = \text{clause}_1 \wedge \text{clause}_2 \wedge \text{clause}_3 \wedge \dots \wedge \text{clause}_k$$

where each clause is an or of literals, with a literal being some x_i or its negation $\neg x_i$.

Note (Disjunctive normal form). For every [conjunctive normal form](#), there is an equivalent way to write it in the so-called *disjunctive normal form*.

Definition A.1.2 (k -CNF). A *k -CNF formula* is a [CNF](#) formula in which each clause has exactly k literals from distinct variables.

Example (3-CNF). A [3-CNF formula](#) can be like

$$\varphi = (x_1 \vee \neg x_2 \vee x_4) \wedge (\neg x_3 \vee x_4 \vee x_5) \wedge (\neg x_1 \vee \neg x_5 \vee \neg x_6).$$

Now, the boolean satisfiability problem asks the following question: given a [\$k\$ -CNF formula](#) φ , does an assignment exist such that φ is evaluated as true? Formally, we have [Problem A.1.1](#).

Problem A.1.1 (k -SAT). Given a [\$k\$ -CNF formula](#) φ , the *k -SAT* problem asks whether φ is satisfiable.

Instead of looking at a general k , we consider a simple but also hard enough case when $k = 3$. Specifically, we ask the following question.

Problem A.1.2 (MAX-3SAT). Given a [3-CNF formula](#) φ and ℓ , the *MAX-3SAT* problem asks is there an assignment of variables such that it satisfies at least ℓ clauses?

Remark. We often call [MAX-3SAT](#) as 3SAT for brevity.

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