

# Assignment 1 - Homework Exercises on Approximation Algorithms

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## A.I-1

We will show that the approximation ratio of the *GreedySchedulingAlgorithm* is at least  $2 - \frac{1}{m}$  by showing an example as follow.

Let's consider this setting:

- 3 machines: M1, M2, M3
- 7 jobs: 1, 1, 1, 2, 1, 1, 2

*GreedySchedulingAlgorithm* will come up with this scheduling:

- M1: 1 2 2
- M2: 1 1
- M3: 1 1

Thus  $ALG = makespan = 1 + 2 + 2 = 5$

We know that:

$$OPT \geq Average_{load} \tag{1}$$

Where  $Average_{load} = \frac{1}{m} \sum_{i=1}^n j_i = \frac{9}{3} = 3$

Here we can find a solution with  $makespan = 3$ . That is:

- M1: 1 2
- M2: 1 2

- M3: 1 1 1

Therefore,  $OPT = 3$

Thus, the approximation ratio is:

$$\rho = \frac{ALG}{OPT} = \frac{5}{3} \quad (2)$$

According to the theorem, the estimated ratio is:

$$\rho_{estimated} = 2 - \frac{1}{m} = 2 - \frac{1}{3} = \frac{5}{3} \quad (3)$$

From 2 and 3, we have  $\rho_{estimated} = \rho$ . Therefore, this bound is tight.

## A.I-2

From the question, we know that

$$\begin{aligned} m &= 10 \\ \sum_{j=1}^n t_j &\geq 1000 \\ t_j &\in [1, 20] ; \text{ for all } i \leq j \leq n \end{aligned}$$

Let  $T'_i$  denote the load of  $M_i$  before  $t_j^*$ , last job, is assigned to the machine. Thus  $T_i^*$ , which represents makespan of the assignment, equals to

$$T_i^* = T'_i + t_j^*$$

Because  $T'_{i^*}$  is the minimum load among all machines, so that we can derive

$$T'_{i^*} \leq \frac{1}{m} \sum_{i=1}^m T'_i = \sum_{j=1}^{j^*} t_j \leq \frac{1}{m} \left[ \sum_{j=1}^n t_j - t_j^* \right] \leq LB$$

Then we can derive

$$\begin{aligned}
T_i^* &= T_{i^*}' + t_j^* \\
&\leq \frac{1}{m} \left[ \sum_{j=1}^n t_j - t_j^* \right] + t_j^* \\
&\leq \frac{1}{m} \sum_{j=1}^n t_j + \left(1 - \frac{1}{m}\right) t_j^* \\
&\leq 100 + \left(1 - \frac{1}{10}\right) 20 \\
&\leq 118
\end{aligned}$$

According *Algorithm Greedy Scheduling* and the question, we know

$$\begin{aligned}
\max \left( \frac{1}{m} \sum_{j=1}^n t_j, \max_{1 \leq j \leq n} (t_j) \right) &\leq LB \leq OPT \\
\max_{1 \leq j \leq n} (t_j) &= 20
\end{aligned}$$

Then we can derive

$$\max \left( \frac{1}{m} \sum_{j=1}^n t_j, 20 \right) \leq LB$$

Since  $\frac{1}{m} \sum_{j=1}^n t_j \geq 1000$ , thus

$$\begin{aligned}
100 &\leq LB \\
&\leq OPT
\end{aligned}$$

Therefore, approximation-ratio( $\rho$ ) equals to

$$\begin{aligned}
T_i^* &\leq \rho OPT \\
\frac{118}{100} &\leq \rho \\
1.18 &\leq \rho
\end{aligned}$$

For this particular setting, *Algorithm Greedy Scheduling* is 1.18 approximation algorithm.

## AI-3-i)

Assume we have the optimal solution, which has  $n$  squares

$$n \leq LB \leq OPT$$

**Lemma 1.** *Each unit square in the grid can overlap at most 4 cells. Let  $n_s$  be the number of square in the integer grid solution. Thus*

$$n_s \leq 4n \leq 4OPT$$

## A.I-3-ii

We propose the algorithm as follow.

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**Algorithm 1** Finding minimum row square cover

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**Require:** Set of Points  $P$

**Ensure:** Min Square Cover  $min$

**Operation:**

```
set currentCoveringPosition = 0
QuickSortAscending( $S$ )
for all Point  $p$  in  $P$  do
  if  $p.x \leq \textit{currentCoveringPosition}$  then
    create square  $s = (p.x, 1, p.x + 1, 0)$ ;
    add  $s$  to  $S$ 
    set currentCoveringPosition =  $p.x + 1$ 
  end if
  set  $min = \textit{sizeof} S$ 
  return  $min$ 
end for
```

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This algorithm is correct because:

- Every point in  $p$  will be covered by a square
- There are no intersections between the squares because we traverse in one direction

This algorithm consists of 2 parts: QuickSort and Traversing the Point to create squares. Let  $t$  be the run time of this algorithm,  $t_{quicksort}$  be the time for quick-sort, and  $t_{assign}$  be the time for creating the squares. We have:

$$t = t_{quicksort} + t_{assign} \leq n \log n + n = O(n \log n) \quad (4)$$

Thus the runtime of this algorithm is  $O(n \log n)$ .

### A.I.3.iii

The idea of our algorithm is that, we put all the points in to a coordinate system, then we divide the coordinate system into a set of unit rows (i.e. rows with height 1). For each row, we use algorithm 1 to find the minimum size square-cover. The global min-square-cover is the sum of all row-square-cover.

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**Algorithm 2** Finding global minimum square cover

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**Input:** Set of points  $P$

**Output:** Min Square Cover  $min$

**Operation:**

```

currentMin = 0;
for all Row  $r$  in the space do
    currentMin += FindRowMinSquare()
end for
set  $min = currentMin$ 
return min

```

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**Theorem.** FindingGlobalMinimumSC is 2 – approximation

*Proof.* We prove this Theorem by induction.

If the optimal solution consists of only 1 square, then  $OPT_1 = 1$ . After applying our algorithm, the square can be split into at most 2 squares.

This is true because if the algorithm returns more than 2 squares, then there is a row which consists of more than 1 square. It means the margin of our points is larger than 1, then the optimal must have more than 1 squares to fit them. It contradicts with our assumption that  $OPT_1 = 1$ .

So  $ALG_1 \leq 2 = 2OPT_1$

Suppose when  $n = k$ , the algorithm is true, that is  $ALG_k \leq 2OPT_k$

Now we add some additional points which insert another square into the optimal solution. Now  $n = k + 1$ .

We have  $OPT_{k+1} = OPT_k + 1$

Applying our algorithm, the final result is the sum of the original input ( $n = k$ ) and the new input ( $n = 1$ ). We know that the Theorem holds for both of them. We have:

$$ALG_{k+1} = ALG_k + ALG_1 \leq 2OPT_k + 2 = 2(OPT_k + 1) = 2OPT_{k+1} \quad (5)$$

Thus the Theorem also holds for  $n = k + 1$ . Therefore, it holds for all values of  $n$ .

In conclusion, this algorithm is 2-*approximation*.

□

## AII.1

(i)

We prove this statement by contradiction.

- Suppose that  $V \setminus C$  is not an independent set of  $G$ . Then there exists a pair of vertices  $(u, v)$  in  $V \setminus C$  which are connected by an edge  $e \in E$ . Thus, both  $u$  and  $v$  are not in  $C$ . Therefore,  $C$  is not the vertex cover of  $G$  anymore.
- Suppose  $C$  is not the vertex cover of  $G$ , then there exists a pair of vertices  $(u, v)$  that are connected by an edge  $e \in E$  but are not in  $C$ . Thus,  $u \in (V \setminus C)$  and  $v \in (V \setminus C)$ . Therefore,  $(V \setminus C)$  is not the vertex cover of  $G$  anymore.

From the reasoning above, we can state that:  $C$  is the vertex cover of  $G$  if and only if  $V \setminus C$  is an independent set of  $G$ .

(ii)

We prove that *ApproxMaxIndependentSet* is not a 2-approximation algorithm by showing a counter example. That is, consider a complete graph, for example, a graph  $G = (V, E)$  where  $V = \{x_1, x_2\}$  and  $E = \{(x_1, x_2)\}$ .

Applying the *ApproxMinVertexCover*( $G$ ), we get  $C = x_1, x_2$  (picking both vertices from the edge).

Now we take the approx max independent set  $ALG = V \setminus C = \emptyset$ .

The optimal solution now is  $OPT = 1$  (picking  $x_1$  or  $x_2$ ).

The approximation ratio is  $\rho = \frac{OPT}{ALG} = \infty \neq 2$ .

So the approximation ratio is not 2.

## AII.2

(i)

The best possible scenario in the presented case is similar to following:

$$(x_1 \vee x_2 \vee x_3) \wedge (x_1 \vee x_4 \vee x_5) \wedge (x_1 \vee x_6 \vee x_7)$$

It can be seen that in this case  $x_1$  is present in all the clauses of the DNF therefore we can eliminate all the clauses of the equation in the first run. This gives us a following lower bound:

$$LB = 1$$

We deduce then:

$$OPT \geq 1$$

We define approximation ratio  $\rho$  as:

$$\rho = \frac{ALG}{OPT}$$

Since we don't know if duplication of elements in a clause is allowed or not we will examine two possible scenarios.

**Duplication allowed:**

Consider the case when:

$$(x_1 \vee x_2 \vee x_2) \wedge (x_1 \vee x_3 \vee x_3) \wedge (x_1 \vee x_4 \vee x_4) \dots \wedge (x_1 \vee x_n \vee x_n)$$

We choose in each iteration elements unique for the clause like  $x_2, x_3, x_4 \dots x_n$ . So we end up having chosen  $n$  elements, excluding the one that was common in all the clauses. That gives us:

$$ALG = (n - 1)$$

$$\rho = (n - 1)$$

**Duplication disallowed:**

If duplication is not allowed as in the following DNF:

$$(x_1 \vee x_2 \vee x_3) \wedge (x_1 \vee x_2 \vee x_4) \wedge (x_1 \vee x_2 \vee x_5) \dots \wedge (x_1 \vee x_2 \vee x_n)$$

We can see that the algorithm chooses  $n - 2$  resulting in :

$$ALG = (n - 2)$$

$$\rho = (n - 2)$$

**(ii)**

For the algorithm to become a 3-approximation algorithm it should be modified so that in each iteration it chooses all three elements in a clause and eliminate all the clauses in DNF that contain any of these three elements.

**Proof**

Let  $D^*$  be a subset of  $D$  that only contains clauses that don't share any variables.

The optimal solution to the problem  $OPT$  contains at least one variable from each clause therefore :

$$OPT \geq |D^*| = LB$$

In our algorithm after the modification, we select three variables in each clause in  $D^*$ , because there are no clauses that share common variables with



them. The other clauses which have common variables as any clause in  $D^*$  are deleted. Thus:

$$ALG \leq 3|D^*| = 3 * LB$$

So it is a 3-approx algorithm.

## AII.3

### (i)

Suppose a  $d$ -hypergraph  $G = (V, E)$  which every edge  $e \in E$  incident to  $d$  vertices in  $V$ . To formulate 0/1 linear programming, we introduce  $X = \{x_1, x_2, \dots, x_n\}$  which  $x_i$  represents  $v_i \in V$  in a linear programming. If  $x_i = 1$ , it means we pick  $v_i$  to the set of double vertex cover,  $C \subset V$ , and otherwise  $x_i = 0$ . For this solution, we want to find a minimum double vertex cover which requires at least 2 vertices from each edge are in  $C$ . Then, we can derive a constraint for 0/1 linear programming

$$\sum_{v_i \in e} x_i \geq 2 \text{ for all } e \in E ;$$

Thus, we then formulate the linear programming.

$$\begin{array}{ll} \text{Minimize} & \sum_{i=1}^n x_i \\ \text{Subject to} & \\ & \sum_{v_i \in e} x_i \geq 2 \quad \text{for all } e \in E ; \text{ at least 2 vertices are selected.} \\ & x_i = \{0,1\} \quad \text{for all } x_i \in X \end{array}$$

### (ii)

Because we can not solve 0/1 linear program in polynomial time, what we have to do next is to relax the program to be a normal linear program by replacing  $\{0,1\}$  constraint with  $0 \leq x \leq 1$

Thus, the linear program is

$$\begin{array}{ll}
\text{Minimize} & \sum_{x=1}^n x_i \\
\\
\text{Subject to} & \\
& \sum_{v_i \in e} x_i \geq 2 \quad \text{for all } e \in E ; \text{ at least 2 vertices are selected.} \\
& 0 \leq x_i \leq 1 \quad \text{for all } x_i \in X
\end{array}$$

Let  $\tau$  denote the rounding threshold such that

$$x_i = \begin{cases} 1, & \text{if } x_i \geq \tau \\ 0, & \text{otherwise} \end{cases}$$

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**Algorithm 3** Finding double vertex cover

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**Input:**  $V, E$

**Output:** A minimum double vertex cover

**Operation:**

Solve the relaxed linear program corresponding to the given problem.

$$\text{Minimize} \quad \sum_{x=1}^n x_i$$

Subject to

$$\sum_{v_i \in e} x_i \geq 2 \quad \text{for all } e \in E$$

$$0 \leq x_i \leq 1 \quad \text{for all } x_i \in X$$

$$C \leftarrow \{ v_i \in V : x_i \geq \tau \}$$

**return** C

---

The next step is to derive  $\tau$  such that all constraints are satisfied and the algorithm always return a valid solution. Let denote  $x^*$  to be an ideal value

of any  $x_i$  such that it satisfies all constraints.

$$\begin{aligned}
\sum_{v_i \in e} x_i &\geq 2 \\
\sum_{i=1}^d x_i &\geq 2 \\
1 + \sum_{i=1}^{d-1} x^i &\geq 2 \\
(d-1)x^i &\geq 1 \\
x_i &\geq \frac{1}{d-1} \\
\therefore \tau &= \frac{1}{d-1}
\end{aligned}$$

Let denote  $W$  denote the value of an optimal to the relaxed linear program and  $OPT$  denote the minium number of double vertex cover. Then  $OPT \geq W$ .

Now we can derive,

$$\begin{aligned}
|C| &= \sum_{v_i \in C} 1 \\
&\leq \sum_{v_i \in C} (d-1)x_i \\
&\leq (d-1) \sum_{v_i \in C} x_i \\
&\leq (d-1)W \\
&\leq (d-1)OPT
\end{aligned}$$

(iii)

Lets take an example of a complete 3-hypergraph, where the optimal double vertex cover is  $|V| - 1$  to make sure every edge has at least 2 vertices selected. So the result of the 0/1-LP is  $|V| - 1$ .

The relaxed-LP formulation is as follow:

- Minimize  $\sum_{i=1}^n x_i$
- Subject to:  $\sum_{x_j \in e} x_j \geq 2$  for all edge  $e$  AND  $0 \leq x_i \leq 1$

We run the algorithm by performing that relaxed-LP on the complete 3-hypergraph, and then round the result following the condition  $x \geq \frac{1}{2}$ .

For the complete 3-hypergraph, the relaxed-LP will return  $x_i = \frac{2}{3}$  for all  $i$  so that each sum of vertices in an edge is 2.

Then the algorithm will pick all of the vertices because they satisfy the condition. The result is:

$$ALG = \frac{2}{3}|V|$$

The integrality gap is:

$$integrality = \frac{|V| - 1}{\frac{2|V|}{3}} = \frac{3}{2} - \frac{3}{2|V|}$$

### AIII-1-i)

We assume total time  $T = \sum_{j=1}^n t_j$  to be the total time of all the jobs. Since we define large job as having time  $t \geq \epsilon T$ , we can deduce that if all running jobs are large jobs, the maximal number of those jobs is equal to :

$$n_{max} = \frac{T}{\epsilon T} = \frac{1}{\epsilon}$$

For each job we take into account there are two possible ways of assigning it to a machine. Therefore the possible ways jobs can be scheduled is:

$$2 * 2 * \dots * 2 = 2^{\frac{1}{\epsilon}}$$

Since we don't distinguish between the machines, we remove the duplicates leaving the total number of schedules at :

$$\frac{2^{\frac{1}{\epsilon}}}{2} = 2^{\frac{1}{\epsilon}-1}$$

## AIII.2

(i)

Let  $d$  denote the distance between 2 arbitrary vertices corresponding to  $P$  and  $d^*$  denote the distance after rounding  $p_{i,x}, p_{i,y}$  where  $p_{i,x}$  and  $p_{i,y}$  denote the x- and y-coordinate of  $p_i \in P$ , by  $\Delta$ .

$$d = \sqrt{(p_{i,x} - p_{j,x})^2 + (p_{i,y} - p_{j,y})^2}$$

$$d^* = \sqrt{(p_{i,x^*} - p_{j,x^*})^2 + (p_{i,y^*} - p_{j,y^*})^2}$$

We know that the range of  $p_x^*$  is

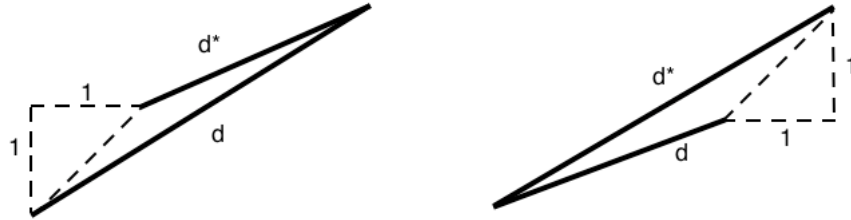
$$\frac{px}{\Delta} \leq p_x^* \leq \frac{px}{\Delta} + 1$$

Then we derive the range of  $d^*$

$$\sqrt{\left(\frac{p_{i,x}}{\Delta} - \left(\frac{p_{j,x}}{\Delta} + 1\right)\right)^2 + \left(\frac{p_{i,y}}{\Delta} - \left(\frac{p_{j,y}}{\Delta} + 1\right)\right)^2} \leq d^* \leq \sqrt{\left(\frac{p_{i,x}}{\Delta} + 1 - \frac{p_{j,x}}{\Delta}\right)^2 + \left(\frac{p_{i,y}}{\Delta} + 1 - \frac{p_{j,y}}{\Delta}\right)^2}$$

From the triangle inequality property, such that a, b and c are the length of the triangle edges

$$c \leq a + b$$



We can simplify the range of  $d^*$  to

$$\sqrt{\left(\frac{p_{i,x}}{\Delta} - \frac{p_{j,x}}{\Delta}\right)^2 + \left(\frac{p_{i,y}}{\Delta} - \frac{p_{j,y}}{\Delta}\right)^2} - \sqrt{2} \leq d^* \leq \sqrt{\left(\frac{p_{i,x}}{\Delta} - \frac{p_{j,x}}{\Delta}\right)^2 + \left(\frac{p_{i,y}}{\Delta} - \frac{p_{j,y}}{\Delta}\right)^2} + \sqrt{2}$$

$$\frac{d}{\Delta} - \sqrt{2} \leq d^* \leq \frac{d}{\Delta} + \sqrt{2}$$

Hence, the error of  $d^*$  is  $2\sqrt{2}$  at most.  
Therefore,

$$2n\sqrt{2}\Delta = \varepsilon OPT$$

$$\Delta = \frac{\varepsilon OPT}{2n\sqrt{2}}$$

(ii)

Let  $P$  and  $P^*$  denote set of edges from the optimal solution and the PTAS algorithm respectively and we know that  $length^*(T) \geq length^*(T^*)$ , then we have

$$\sum_{p_i, p_j \in P^*} d_{ij}^* \leq \sum_{p_i, p_j \in P} d_{ij}^*$$

Thus, we can derive

$$\begin{aligned}
length(T^*) &= \sum_{p_i, p_j \in P^*} d_{ij} \\
&\leq \sum_{p_i, p_j \in P^*} \Delta(d_{ij}^* + \sqrt{2}) \\
&\leq \Delta \sum_{p_i, p_j \in P^*} (d_{ij}^* + \sqrt{2}) \\
&\leq \Delta \sum_{p_i, p_j \in P} (d_{ij}^* + \sqrt{2}) \\
&\leq \Delta \sum_{p_i, p_j \in P} \left( \frac{d_{ij}}{\Delta} + 2\sqrt{2} \right) \\
&\leq \sum_{p_i, p_j \in P} d_{ij} + \Delta \sum_{p_i, p_j \in P} 2\sqrt{2} \\
&\leq length(T) + \Delta 2|P|\sqrt{2} \\
&\leq OPT + \left( \frac{\varepsilon OPT}{2n\sqrt{2}} \right) 2n\sqrt{2} \\
&\leq (1 + \epsilon)OPT
\end{aligned}$$

(iii)

Let  $m^*$  denote the new boundary of the coordinate after rounding  $p_x, p_y$  to  $p_x^*, p_y^*$  and we also know that

$$\begin{aligned}
m &= \max(p_x, p_y) \\
OPT &\geq 2m
\end{aligned}$$

Thus

$$\frac{m}{\Delta} \leq m^* \leq \frac{m}{\Delta} + 1$$

Then, we can derive the running time

$$\begin{aligned}
m^* &\leq \frac{m}{\Delta} + 1 \\
&\leq \frac{m2n\sqrt{2}}{\epsilon OPT} + 1 \\
&\leq \frac{m2n\sqrt{2}}{\epsilon 2m} + 1 \\
&\leq \frac{n\sqrt{2}}{\epsilon} + 1
\end{aligned}$$

Therefore, the running time is

$$\begin{aligned}
O(nm^*) &= O\left(n \frac{n\sqrt{2}}{\epsilon} + 1\right) \\
&= O\left(\frac{n^2\sqrt{2}}{\epsilon}\right)
\end{aligned}$$

### AIII.3

(i)

Because we know that  $ALG(G, \epsilon) \in \mathbb{N}$ , so that if we can find such  $\epsilon$  that the algorithm yields

$$OPT - 1 < ALG(G, \epsilon) \leq OPT$$

Then, we can get  $OPT$  in polynomial time.

In order to get such  $\epsilon$ , we will derive

$$\begin{aligned}
ALG(G, \epsilon) &> OPT - 1 \\
&> \left(1 - \frac{1}{OPT}\right) OPT
\end{aligned}$$

Hence we can get  $OPT$  if we choose  $\epsilon < \frac{1}{OPT}$  and we also know that the algorithm uses  $ALG(G, \epsilon)$  as a subroutine.

Therefore, there is no such FPTAS exist.



**(ii)**

The proof above indeed implies that there is no PTAS such a problem because we know that a PTAS algorithm also computes a  $(1 - \epsilon)$ -approximation for the problem and if we choose  $\epsilon > \frac{1}{OPT}$  as the proof above then, the PTAS algorithm will yield  $OPT$  in polynomial time of  $n$ .

Therefore there is no PTAS exist anymore.