

SAPIENZA UNIVERSITY OF ROME

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# Algorithm Design Second Homework

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Andrea Fioraldi 1692419

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## EXERCISE 1

We redefine our problem as a problem on graphs. Let  $G(M, F, E)$  a fully connected graph in which each friend is a node. Each edge  $(u, v)$  has a cost  $w(u, v) \geq 0$ . We define the following functions:

$$\begin{aligned} \deg(v, G) &:= \sum_{u \in G.M \cup G.F} w(v, u); \\ \text{density}(G) &:= \frac{1}{|G.M \cup G.F|} \sum_{v, u \in G.M \cup G.F} w(v, u); \end{aligned}$$

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1 algorithm michele_party_approx( $G$ )
2    $n := |G.M \cup G.F|$ 
3    $H_n := G$ 
4    $\text{max\_d} := -1$ 
5    $s := -1$ 
6   for  $i = \frac{n}{2}$  to 1
7      $d := \text{density}(H_i)$ 
8     if  $d > \text{max\_d}$ 
9        $\text{max\_d} = d$ 
10       $s = i$ 
11       $m := v \in H_i.M \mid \deg(v, H_i) \leq \deg(u, H_i) \ \forall u \in H_i.M$ 
12       $f := v \in H_i.F \mid \deg(v, H_i) \leq \deg(u, H_i) \ \forall u \in H_i.F$ 
13       $H_{i-1} = H_i \setminus \{m, f\}$ 
14 return  $H_s$ 

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Let  $OPT$  the optimal solution and  $d_{OPT} = \frac{1}{|OPT|} \sum_{v, u \in OPT} w(v, u) = \frac{W}{N}$  its density. Let  $m \in OPT \cap M$  and  $f \in OPT \cap F$ . Consider that  $\frac{W}{N} \geq \frac{W - \deg(m) - \deg(f)}{N-2} \Leftrightarrow \frac{\deg(m) + \deg(f)}{N-2} \geq \frac{W}{N-2} - \frac{W}{N} \Leftrightarrow \deg(m) + \deg(f) \geq W - \frac{N-2}{N} * W = 2 * \frac{W}{N}$ . So in  $OPT$  the sum of the degree of a generic  $m$  and  $f$  is at least twice  $d_{OPT}$ .

Consider the iteration  $i$  of the algorithm in which for the first time two nodes  $m \in OPT \cap M$  and  $f \in OPT \cap F$  are removed. In the  $H_i$  graph all the pairs  $m \in H_i \cap M$  and  $f \in H_i \cap F$  have  $\deg(m) + \deg(f) \geq 2 * d_{OPT}$  thanks to the fact that we are going to remove nodes (so with minimal sum of degrees) in which the previous observation holds. In  $H_i$  there are  $\frac{|H_i|}{2}$  pairs that will be removed until  $i = 1$ . These couples do not share nodes. The total cost of the edges in  $H_i$  is greater than  $\left(2 * d_{OPT} * \frac{|H_i|}{2}\right) / 2$  (the division by 2 is due to the fact that we want to avoid to consider edges twice). So we have that the density is greater than  $\frac{2 * d_{OPT} * \frac{|H_i|}{2}}{2 * |H_i|} = \frac{d_{OPT}}{2}$ . Since the algorithm returns the graph with the highest density over all the iterations we have a solution with density at least  $\frac{d_{OPT}}{2}$ . We proved that *michele\_party\_approx* is a 2-approximation. The cost is  $O(n^3)$  because the most expensive operation in the body of the loop is the density that cost  $|E| = n^2$ .

## EXERCISE 2

Our solution was inspired by the Set Cover approximation proof described in [1] but follows the path of the proof shown during the class.

Let  $A$  is the set of required skills.  $S$  is the set of all the people available, each people is represented as a set of skills  $S_j \subseteq A$ . Let  $n = |A|$ .

We can express the Set Cover with Redundancies problem using the following ILP formulation:

$$\begin{aligned} \min \quad & \sum_{S_j \in S} c_j * x_j \quad \text{s.t.} \\ & \sum_{S_j | A_i \in S_j} x_j \geq 3 \quad \forall A_i \in A \\ & x_j \in \{0, 1\} \quad \forall S_j \in S \end{aligned}$$

In order to build a randomized approximation consider the associated LP relaxation where  $x_j^* \in [0, 1]$ . The LP solution is a vector  $x^*$  of real values.

Consider the algorithm ALG in which each person  $S_j$  is chosen randomly with probability  $p_j = \min(d * \log(n) * x_j^*, 1)$  with all choices that are independent. We denote the vector of these choices as  $x'$ .

Let  $C_i = \sum_{S_j | A_i \in S_j} x'_j$  a random variable that represents the times that the skill  $A_i$  is covered. The expectation of  $C_i$  is  $\mathbb{E}[C_i] = \mathbb{E}\left[\sum_{S_j | A_i \in S_j} x'_j\right] = \sum_{S_j | A_i \in S_j} p_j$ . Then,  $\mathbb{E}[C_i] = d * \log(n) * \sum_{S_j | A_i \in S_j} x_j^* \geq d * \log(n) * 3$  because the cover constraint in LP is satisfied.

From this, thanks to the Chernoff lower bound, follows that:

$$\Pr[C_i < 3] = \Pr\left[C_i < \left(1 - \left(1 - \frac{1}{d * \log(n)}\right)\right) * d * \log(n) * 3\right] \leq \Pr\left[C_i < \left(1 - \left(1 - \frac{1}{d}\right)\right) * d * \log(n) * 3\right] \leq \exp\left(-\frac{1}{2} * 3 * d * \log(n) * \left(1 - \frac{1}{d}\right)^2\right)$$

Note that  $\left(1 - \frac{d * \log(n) - 1}{d * \log(n)}\right) * d * \log(n) = 1$ .  $0 < 1 - 1/d < 1$  must be valid in order apply Chernoff and so we have that  $d > 1$ .

We have that the probability that the skill  $A_i$  is covered more than 3 times is  $\Pr[C_i \geq 3] \geq 1 - \exp\left(-\frac{1}{2} * 3 * d * \log(n) * \left(1 - \frac{1}{d}\right)^2\right)$ .

The probability that at least one  $A_i$  is not covered 3 times is  $\Pr\left[\bigcup_{A_i \in A} C_i < 3\right] \leq$

$$\sum_{A_i \in A} \Pr[C_i < 3] = n * \exp\left(-\frac{1}{2} * 3 * d * \log(n) * \left(1 - \frac{1}{d}\right)^2\right).$$

Note that increasing  $d$  give us a better probability that all skills are covered but also, on the other hand, decrease the probability to have a minimal solution because over a threshold we have that all the variables in  $x'$  are 1.

The expected cost is  $\mathbb{E}\left[\sum_{S_j \in S} c_j * x'_j\right] = \sum_{S_j \in S} c_j * \mathbb{E}[x'_j] = \sum_{S_j \in S} c_j * d * \log(n) * x_j^*$  and so it is  $d * \log(n)$  times the cost of the LP.

With the Markov inequality, we bound with a parameter  $k$  the probability to fail on having an average cost of less than  $k * d * \log(n)$  times the cost of LP.

$$Pr \left[ \sum_{S_j \in S} c_j * x'_j \geq k * d * \log(n) * \sum_{S_j \in S} c_j * x_j^* \right] \leq \frac{\sum_{S_j \in S} d * \log(n) * c_j * x_j^*}{\sum_{S_j \in S} k * d * \log(n) * c_j * x_j^*} = \frac{1}{k}$$

Obviously this is an event that we want to avoid and setting a big  $k$  help us in that but, on the contrary, give us also a worse approximation factor than  $d * \log(n)$ .

We must choose  $d$  and  $k$  in a way that maximizes the probability that our solution is valid and the expected cost is in the bound and maximize the probability that this solution is minimal.

Choosing  $d = 3$  we have an interesting result:

$$Pr \left[ \bigcup_{A_i \in A} C_i < 3 \right] = n * \exp(-2 * \log(n)) = 1/n$$

The probability that all the skills are covered, so that  $x'$  is feasible, is  $1 - 1/n$  that is quite high and so we expect to need only a run to get a feasible solution.

The cost of this approximation must be at most  $3 * k * \log(n) * LP$ .

The probability that the solution is not feasible or that the cost exceed the bound is

$$Pr \left[ \mathbb{E} \left[ \sum_{S_j \in S} c_j * x'_j \right] \geq \sum_{S_j \in S} k * d * \log(n) * x_j^* \vee \bigcup_{A_i \in A} C_i < 3 \right] \leq \frac{1}{k} + \frac{1}{n}.$$

Setting  $k = 2$  we get that the solution is feasible and with cost less than  $6 * \log(n) * LP$  with probability greater than  $\frac{1}{2} - \frac{1}{2 * n}$  that is almost  $\frac{1}{2}$  with a reasonable  $n$ . We conclude that, as in the normal set cover approximation, the expected number of repetitions are 2 and we choose  $d$  and  $k$  with the intention to have this a number of repetitions.

### EXERCISE 3

We denote as  $F^*$  the optimal solution of the problem.  $F^*$  is the minimum cost set of edges that if removed creates  $k$  connected components with each target terminal  $s_i$  inside each component.

Let  $F_i^* \subset F^*$  the set of edges that if removed separated the vertex  $s_i$  from the other  $s_j$  with  $i \neq j$ . We have that  $\cup_{i=0}^k F_i^* = F^*$ .

Each edge  $e$  in  $F^*$  is contained in two  $F_i^*$  because it is incident at two connected components. So we can say that:

$$\sum_{i=0}^k \sum_{e \in F_i^*} w(e) = 2 * \sum_{e \in F^*} w(e)$$

Our algorithm returns  $k$  min cuts  $F_i$ . By definition  $F_i$  is the minimum set of edges that separates  $s_i$  from the other  $s_j$  with  $i \neq j$  so we have that  $\sum_{e \in F_i} w(e) \leq \sum_{e \in F_i^*} w(e)$ .

This implies that:

$$\sum_{i=0}^k \sum_{e \in F_i} w(e) \leq 2 * \sum_{e \in F^*} w(e)$$

And so, in the worst case, the cost of  $F$  is twice the cost of  $F^*$ . We proved that this is a 2-approximation algorithm.

## EXERCISE 4

Let  $n = |D|$ . Let  $sub(D, i, j)$  the substring of  $D$  from  $i$  to  $j - 1$ . As described in [2] we define a directed graph with  $V = \{0, \dots, n\}$  (the set of the indexes of the nucleobases in  $D$ ) and  $E = \{(i, j) \subset V^2 | i < j \wedge sub(D, i, j) \in G\}$ . A path from node 0 to  $n$  in this graph is a concatenation of genes that results in  $D$ .  $\delta^-(v)$  is the set of incoming edges to node  $v$  and  $\delta^+(v)$  the outgoing edges.

The variables  $x_k$  represents if a gene  $G_k \in G$  is in  $D$ . The variables  $z_{i,j}$  represents if an edge  $(i, j)$  is in the optimal path.

The ILP formulation is the following:

$$\begin{aligned}
 & \min \sum_{k=1}^m w_k * x_k && \text{s.t.} \\
 (1) \quad & x_k |_{G_k=sub(D,i,j)} - z_{i,j} \geq 0 && \forall (i, j) \in E \\
 (2) \quad & \sum_{(i,j) \in \delta^+(0)} z_{i,j} = 1 \\
 (3) \quad & \sum_{(i,j) \in \delta^-(v)} z_{i,j} - \sum_{(i,j) \in \delta^+(v)} z_{i,j} = 0 \quad \forall v \in V \setminus \{0, n\} \\
 & x_k, z_{i,j} \in \{0, 1\}
 \end{aligned}$$

(1) is to force to consider a gene if a substring equal to the gene is taken. (2) and (3) are for the flow integrity.

Consider now the LP-relaxion with  $x_k, z_{i,j} \geq 0$ . Let's compute the dual of it. We associate varibales and constraints as follows: (1)  $\rightarrow q_{i,j}$ , (2)  $\rightarrow p_0$ , (3)  $\rightarrow p_v$  with  $v \in V \setminus \{0, n\}$ ,  $x_k \rightarrow (4)$ ,  $z_{i,j} \rightarrow (5)$ . The  $A$  matrix of the primal has dimension  $(n + |E|) \times (m + |E|)$ . Considering the fact that and edge  $(i, j)$  can be only in a  $\delta^-(v)$  and only in a  $\delta^+(v)$  we can easily compute  $A^T$ . Only constraint (2) contribute to the objective function and only  $q$  is sign constrained (cfr. [3]).

$$\begin{aligned}
 & \max p_0 && \text{s.t.} \\
 (4) \quad & \sum_{i,j | sub(D,i,j)=G_k} q_{i,j} \leq w_k && \forall k \in \{1, \dots, m\} \\
 (5) \quad & -q_{i,j} + p_i - p_j \leq 0 && \forall (i, j) \in E \\
 & q_{i,j} \geq 0
 \end{aligned}$$

## EXERCISE 5

The proposed problem looks like a finite zero-sum game theory problem.

| Comet, Dasher | Head  | Tail  |
|---------------|-------|-------|
| Head          | 4, -4 | -1, 1 |
| Tail          | -2, 2 | 2, -2 |

We create the Comet's payoff matrix  $A_{2,2}$  taking the first value of each table. The matrix  $B_{2,2} = -A$ , on the other hand, is the payoff matrix for Dasher. Santa must assign the probability to get head or tail for each coin in order to convince both Comet and Dasher to play. To do that the game must not be balanced against either player and so we equal the expected payoff of the two players in order to make the game fair.

$$Pr[\text{Comet coins is head}] = p_h \quad Pr[\text{Comet coins is tail}] = p_t \quad p = \begin{pmatrix} p_h \\ p_t \end{pmatrix} = \begin{pmatrix} p_h \\ 1 - p_h \end{pmatrix}$$

$$Pr[\text{Dasher coins is head}] = q_h \quad Pr[\text{Dasher coins is tail}] = q_t \quad q = \begin{pmatrix} q_h \\ q_t \end{pmatrix} = \begin{pmatrix} q_h \\ 1 - q_h \end{pmatrix}$$

The expected payoff of Comet is  $p^T * A * q$  and the expected payoff of Dasher is  $p^T * B * q = -p^T * A * q$ . When they are equal we have that  $2 * p^T * A * q = 0$  and so  $p^T * A * q = 0$  is out constraint in order to have a fair game.

$$(p_h \quad 1 - p_h) * \begin{pmatrix} 4 & -1 \\ -2 & 2 \end{pmatrix} * \begin{pmatrix} q_h \\ 1 - q_h \end{pmatrix} = (p_h \quad 1 - p_h) * \begin{pmatrix} 5 * q_h - 1 \\ -4 * q_h + 2 \end{pmatrix} = p_h * (5 * q_h - 1) + (1 - p_h) * (-4 * q_h + 2) = 9 * p_h * q_h - 3 * p_h - 4 * q_h + 2$$

$9 * p_h * q_h - 3 * p_h - 4 * q_h + 2 = 0$  with  $p_h, q_h \in [0, 1]$  is the constraint that give to santa a method how to select the probabilities in order to have a fair game.

Note that this equation is a piece of an hyperbola. A valid solution is, as instance,  $p_h = 1 \wedge p_t = 0 \wedge q_h = \frac{1}{5} \wedge q_t = \frac{4}{5}$ .

## EXERCISE 6

Let  $h$  the position of Giorgio's home. Let  $S(i) = Pr[Safe|x(t) = i]$  the probability that Giorgio goes to home safely when he is at position  $i$ . We know that  $Pr[x(t+1) = x(t) + 1] = p$  and  $Pr[x(t+1) = x(t) - 1] = 1 - p$ . Follows that:

$$S(i) = \begin{cases} 0 & \text{if } i = -1 \\ 1 & \text{if } i = h \\ p * S(i-1) + (1-p) * S(i+1) & \text{otherwise} \end{cases}$$

We have that in general  $S(i) = p * S(i-1) + (1-p) * S(i+1)$  and  $S(i) = p * S(i) + (1-p) * S(i)$  (from  $1 = p - (1-p)$ ) and so  $S(i+1) = \frac{p}{1-p} * (S(i) - S(i-1)) + S(i)$ . Follows that  $S(i+2) = \frac{p}{1-p} * (S(i+1) - S(i)) + S(i+1) = \frac{p}{1-p} * \left( \frac{p}{1-p} (S(i) - S(i-1)) + S(i) - S(i) \right) + \frac{p}{1-p} * (S(i) - S(i-1)) + S(i)$ .

With  $i = 0$  we have  $S(2) = \left(\frac{p}{1-p}\right)^2 * S(0) + \frac{p}{1-p} * S(0) + S(0)$  and so clearly:

$$S(i) = \sum_{j=0}^i \left( \frac{p}{1-p} \right)^j * S(0).$$

Consider now when  $i = h$ . The problem says that Giorgio makes an infinite number of steps so we can set  $h = +\infty$ .

$$1 = S(+\infty) = S(0) * \sum_{j=0}^{+\infty} \left( \frac{p}{1-p} \right)^j.$$

This is a geometric series [4]. By definition if  $|\frac{p}{1-p}| < 1$  it converges to  $\frac{1}{1-\frac{p}{1-p}}$ .

The probability that Giorgio goes to the hospital from position 0 is  $Pr[Hospital|x(t) = 0] = 1 + Pr[Safe|x(t) = 0] = 1 - S(0)$ .

$|\frac{p}{1-p}| < 1 \Leftrightarrow p < 1/2$  and so, finally, we have that:

$$Pr[Hospital|x(t) = 0] = 1 - \frac{1}{\sum_{j=0}^{+\infty} \left( \frac{p}{1-p} \right)^j} = \begin{cases} \frac{p}{1-p} & p < 1/2 \\ 1 & p \geq 1/2 \end{cases}$$

Giorgio goes to hospital for sure when  $p \geq \frac{1}{2}$  and goes to hospital with probability at most  $\frac{1}{2}$  when  $\frac{p}{1-p} \leq \frac{1}{2} \Rightarrow p \leq \frac{1}{3}$ .



## REFERENCES

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