601.435/635 Approximation Algorithms **Lecturer:** Michael Dinitz
Topic: LP Solutions as Metrics: Multicut **Date:** 3/28/24
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18.1 Introduction

Today we're going to talk about a cut problem known as MULTICUT which is even more general than MULTIWAY CUT. We very briefly discussed this at the end of last lecture, where I described how to use techniques that we already know to design an $O(\log n)$-approximation. While I did this very quickly, all the details are in the lecture notes from last time. Today we're going to improve this slightly to give an $O(\log k)$-approximation.

18.2 Definition and Relaxation

Definition 18.2.1 *In the Multicut problem, we are given a graph $G = (V, E)$ with costs $c : E \to \mathbb{R}^+$, and k pairs $(s_1, t_1), \ldots, (s_k, t_k)$ of nodes. A feasible solution is a set $F \subseteq E$ such that s_i and t_i are not connected in $G \setminus F$ for all $i \in [k]$. The objective is to minimize $\sum_{e \in F} c(e)$.*

For the remainder of the day, we're going to prove the following theorem:

Theorem 18.2.2 *There is an $O(\log k)$-approximation algorithm for Multicut.*

We will use $\mathcal{P}_i$ to denote the set of all s_i-t_i paths. The problem admits the following LP relaxation:

$$\begin{aligned}
\text{minimize:} \quad & \sum_{e \in E} c(e) \cdot x_e & & \textbf{(MULTICUT-LP)} \\
\text{subject to:} \quad & \sum_{e \in P} x_e \geq 1 \quad \forall\, i \in [k], \forall\, P \in \mathcal{P}_i & & (18.2.1) \\
& 0 \leq x_e \leq 1 \quad \text{for each edge } e \in E & & (18.2.2)
\end{aligned}$$

Note: As with multiway cut, we can solve this LP in polytime via ellipsoid, using shortest path (for each $\mathcal{P}_i$) to separate. For the remainder, we will use x to refer to the solution of the LP, and set $V^* = \sum_{e \in E} c(e) x_e$ as the value of the solution.

Definition 18.2.3 *Let $d : V \times V \to \mathbb{R}^+$ be the shortest path metric using the LP solution $\vec{x}$ for the edge lengths.*

Definition 18.2.4 *For all $S \subseteq V$, let $\delta(S) = E(S, \bar{S})$ denote all edges that have exactly one endpoint in S.*

Definition 18.2.5 *For all sets of edges $E' \subseteq E$, let $c(E') = \sum_{e \in E'} c(e)$.*

18.3 Rounding

To move forward, we're going to take inspiration from a physical metaphor. This of each edge as a "pipe". We're thinking of x_e as the length of e, so if we think of $c(e)$ as the "cross-sectional area", then the "volume" of an edge would be $c(e)x_e$. This motivates the following definition.

Definition 18.3.1 (Volume)

$$V(s_i, r) = \frac{V^*}{k} + \sum_{\substack{e=\{u,v\}\in E \\ u,v\in B(s_i,r)}} c(e)x_e + \sum_{\substack{e=\{u,v\}\in E \\ u\in B(s_i,r) \\ v\notin B(s_i,r)}} c(e)(r - d(s_i, u))$$

The second term above should be thought of as the volume of all edge-pipes fully inside the ball around s_i, and the third as (a lower bound for) the volume contained in $B_{G'}(s_i, r)$ of edge-pipes leaving the ball. The first term is included to make later calculations easier.

The next lemma is the main technical piece.

Lemma 18.3.2 (Region-Growing Lemma) *For all $i \in [k]$, we can find in polytime a value $0 \leq r < \frac{1}{2}$ such that:*

$$c(\delta(B(s_i, r))) \leq 2\ln(k+1) \cdot V(s_i, r)$$

Before we prove this lemma, let's show how to approximate MULTICUT if we assume that it is true.

Algorithm 1 Constructing an integer solution

```
Init: F = ∅
for i = 1 to k do
  if s_i, t_i connected in G then
    Let r_i ∈ [0, 1/2) be the r value from the region-growing lemma.
    F ← F ∪ δ(B(s_i, r_i))
    Remove B(s_i, r_i) and all incident edges from the graph.
  end if
end for
return F
```

One important note to clarify this, since we're changing the graph throughout this algorithm: distances, balls and volumes are with respect to the *current* graph, not the original.

Theorem 18.3.3 *The output F from Algorithm 1 is feasible.*

Proof: The only way this might not be feasible is if some $s_i - t_i$ pair are both in $B(s_j, r)$ for some j. But this cannot happen since $r < 1/2$ and $d(s_i, t_i) \geq 1$ throughout the algorithm. ■

Theorem 18.3.4 $c(F) \leq 4\ln(k+1)V^* \leq 4\ln(k+1) \cdot OPT$.

Proof: Let's do some definitions.

- Let B_i be $B(s_i, r_i)$ in iteration i (if we did not create such a ball in iteration i because s_i and t_i were already separated, let $B_i = \emptyset$). Note that since the algorithm changes the graph throughout the algorithm, this might not have been $B(s_i, r_i)$ at the beginning of the algorithm.

- Similarly, let $F_i = \delta(B(s_i, r_i))$ be the edges removed by the algorithm in iteration i. Then clearly $F = \cup_{i=1}^k F_i$, and $F_i \cap F_j = \emptyset$ for all $i \neq j$.

- Let $V_i = \sum_{e=\{u,v\}:u,v\in B_i} c(e)x_e + \sum_{e\in F_i} c(e)x_e$ be the total volume of edges removed in iteration i. Note that $V_i \geq V(s_i, r_i) - \frac{V^*}{k}$, since V_i contains the full volume of edges in F_i while $V(s_i, r_i)$ contains only part of their volume (but with an additional V^*/k).

  Moreover, every edge contributes to V_i for at most one value of i, since the first time at least one of the endpoints is in B_i, the edge is removed from the graph. Thus $\sum_{i=1}^k V_i \leq V^*$

Now note that every edge in F is in exactly one F_i by our definition of the F_i's, and moreover the value r_i was chosen from the region growing lemma. Thus we get that

$$
\begin{aligned}
c(F) = \sum_{i=1}^k c(F_i) &\leq (2\ln(k+1)) \sum_{i=1}^k V(s_i, r_i) \\
&\leq (2\ln(k+1)) \sum_{k=1}^k \left(V_i + \frac{V^*}{k}\right) \\
&\leq 4\ln(k+1) \cdot V^*,
\end{aligned}
$$

as claimed. ■

So now it only remains to prove the Region Growing Lemma (Lemma 18.3.2). For the rest of today, let $c(r) = c(\delta(B(s_i, r))$ and let $V(r) = V(s_i, r)$.

Proof of Lemma 18.3.2: We're eventually going to get a deterministic algorithm, but let's start with a randomized algorithm: choose r uniformly at random from $[0, 1/2)$. We want to show that if we do this, then $E\left[\frac{c(r)}{V(r)}\right] \leq 2\ln(k+1)$.

Order $B(s_i, \frac{1}{2})$ as $\{v_1, \ldots, v_m\}$, where $r_j = d(s_i, v_j)$, and $0 \leq r_1 \leq r_2 \leq \cdots \leq r_m < \frac{1}{2}$. We also define $r_0 = 0$ for later calculations.

Surprisingly, we're going to do a bunch of calculus to prove this. I'm going to abuse calculus a bit here – see the book for the more formally correct way of doing this. Consider the function $V(r)$, which (just to recall) is

$$
V(r) = \frac{V^*}{k} + \sum_{\substack{e=\{u,v\}\in E \\ u,v\in B(s_i,r)}} c(e)x_e + \sum_{\substack{e=\{u,v\}\in E) \\ u\in B(s_i,r) \\ v\notin B(s_i,r)}} c(e)(r - d(s_i, u)).
$$

Unfortunately, $V(r)$ is not continuous or differentiable, since there can be discontinuities at the values $\{r_j\}$. But let's *pretend* like it's differentiable. Note that for $r \in (r_j, r_{j+1})$, for any j, it is in fact differentiable with derivative $\frac{d}{dr}V(r) = c(r)$. This is because the first and second terms are constant in this range of r, so we just need to care about the third term, which gives exactly $c(r)$.

Now we can use calculus to figure out the "average" value of $\frac{c(r)}{V(r)}$ over $[0, \frac{1}{2})$:

$$\begin{aligned}
\frac{1}{1/2}\int_0^{1/2} \frac{c(r)}{V(r)}\,dr &= 2\int_0^{1/2} \frac{1}{V(r)} \cdot \frac{dV(r)}{dr}\,dr \\
&= 2\int_0^{1/2} \frac{1}{V(r)}\,dV(r) \\
&= 2(\ln(V(\tfrac{1}{2})) - \ln(V(0))) \\
&= 2\ln\left(\frac{V(1/2)}{V(0)}\right) \\
&\le 2\ln\left(\frac{V^*/k + V^*}{V^*/k}\right) = 2\ln(k+1)
\end{aligned}$$

It then would follow from the mean value theorem that there exists some $r \in [0, \frac{1}{2})$ achieving the average value. For this r we would then have $\frac{c(r)}{V(r)} \le 2\ln(k+1)$, so that $c(r) \le 2\ln(k+1)V(r)$ as desired.

The analysis above was based on the (false) assumption that $V(r)$ is continuous and differentiable. We will now complete the argument by discarding this assumption. In particular, note that $V(r)$ is piecewise linear and monotone increasing with discontinuities at the r_j's listed above. Then the real average value of $\frac{c(r)}{V(r)}$ over $[0, \frac{1}{2})$ is given by (with r_j^- infinitesimally smaller than r_j):

$$\begin{aligned}
\frac{1}{1/2}\sum_{j=0}^{m}\int_{r_j}^{r_{j+1}^-} \frac{c(r)}{V(r)}\,dr &= 2\sum_{j=0}^{m}\int_{r_j}^{r_{j+1}^-} \frac{1}{V(r)}\,dV(r) \\
&= 2\sum_{j=0}^{m}(\ln(V(r_{j+1}^-) - \ln(V(r_j))) \\
\text{($V(r)$ increasing)} \quad &\le 2\sum_{j=0}^{m}(\ln(V(r_{j+1}) - \ln(V(r_j))) \\
\text{(sum telescopes)} \quad &\le 2(\ln(V(r_m)) - \ln(V(r_0))) \\
(r_0 = 0, r_m \le \tfrac{1}{2}\text{; recall `pretend' section)} \quad &\le 2\ln(k+1)
\end{aligned}$$

Before, we concluded by saying that the MVT allowed us to find an r achieving the average value. Here, because $V(r)$ is increasing and $c(r)$ is constant over each $[r_j, r_{j+1})$ interval, we can say that the smallest value of $\frac{c(r)}{V(r)}$ will occur at some r_j^-. By the above, for $r = r_j^-$ we will then have that $c(r) \le 2\ln(k+1)V(r)$, as desired. And note that there are only $m \le n$ different values of r_j, so we can just check each one and deterministically find the best. ∎