CSC373H1 Summer 2014 Assignment 3

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Q1. Construction:

To construct a family of graphs where G(V, E) has exponential number of minimum cuts between source and terminal, we being with n being the number of vertices that appear between the source and the terminal, we will also assume that the capacity on each edge is 1. If n = 0 then there are no vertices between the source and the terminal and thus V = {s, t} and E = $\emptyset$, Clearly, a single minimum-cut exists since we can only s $\in$ S and t $\in$ T, no edges exist to place a capacity on, so indeed $2^{|V|-2} = 2^0 = 1$.

For arbitrary n, n $\geqslant$ 1, and n = |V| - 2, then G(V, E) would be constructed such that V = {s, t, v_1, v_2, ..., v_n} and E = {(s, v_1), (v_1, t), (s, v_2), (v_2, t), ..., (s, v_n), (v_n, t)} (i.e. the only edges connected to a vertex are those that have s as their startpoint, or t as their endpoint). This arrangement is chosen because every cut in G(V, E) will be a minimum cut. By the arrangement of G(V, E), and since all capacities are the same. Whether a vertex v is or is not in the set S will not affect the summation of outgoing edge capacities, either the capacity from (v, t) or from (s, v), respectively, will be used in computing the minimum cut. Therefore we have a maximum number of minimum cuts in G.

Proof of exponential minimum cuts:

From the description above, in our construction of G(V, E), for any $v_i \in V$ the vertex can be contained within or outside the set that describes a cut in G and the cut is still a minimum cut. If we think of each minimum-cut as being described by a binary string for n = i, i $\geqslant$ 1, (q_{v_1}, q_{v_2}, ..., q_{v_n}), such that:

$$q_{v_i} = \begin{cases} 1 & \text{if } v_i \in \text{minimum cut set} \\ 0 & \text{if } v_i \notin \text{minimum cut set} \end{cases}$$

<u>Base:</u> Let n = 1, and let the B_1 be the set of binary strings describing the minimum cuts in G(V, E). Then B_1 contains two binary strings such that B_1 = {0, 1}. Indeed, $|B_1| = 2 = 2^n = 2^{|V|-2}$.

<u>IH:</u> Let n = i, i > 1, and let B_i be the set of binary strings describing the minimum cuts in G(V, E), such that V = {s, t, v_1, v_2, ..., v_i} and E = {(s, v_1), (v_1, t), (s, v_2), (v_2, t), ..., (s, v_i), (v_i, t)}. Suppose that $|B_i| = 2^i = 2^{|V|-2}$.

<u>IS:</u> Let n = i + 1, it follows from the construction description that V' = {s, t, v_1, v_2, ..., v_i, v_{i+1}} and E' = {(s, v_1), (v_1, t), (s, v_2), (v_2, t), ..., (s, v_i), (v_i, t), (s, v_{i+1}), (v_{i+1}, t)}. Clearly, the addition of v_{i+1} extends G(V, E) to G(V', E'). The minimum cuts in G(V', E') are composed from the minimum cuts in G(V, E) except that v_{i+1} is either inside or outside these previous cuts. Thus, to compute the set of minimum cuts as binary strings, B_{i+1}, simply take each binary string in B_i and append a 0 to be a minimum cut where v_{i+1} is outside the set, and then again take each binary string in B_i and append a 1 to be a minimum cut where v_{i+1} is inside the set. Thus, $|B_{i+1}| = 2 * (|B_i|) = 2 * 2^i = 2^{i+1}$.

By proof of induction this construction will have exponential minimum cuts between the source and the terminal. ■

Q2. Claim:

Function f, such that f(S) is the number of edges (u, v) with $u \in S$, $v \in V \backslash S$, is submodular.

Proof:

Using the fact provided in the question, all that is necessary to show that f is submodular is to show that for any two subsets $A, B \subseteq V$, $f(A) + f(B) \geqslant f(A \cup B) + f(A \cap B)$.

Consider the set of edges in G(V, E), $E = \{e_1, e_2, ..., e_n\}$. If we think of f(S) as an iterative algorithm that goes through each $e_i \in E$ and checks whether the startpoint of $e_i \in S$ and whether the endpoint of $e_i \in V \backslash S$. Then we have $f_{e_i}(S)$:

$$f_{e_i}(S) = \begin{cases} 1 & \text{if startpoint of } e_i \text{ in S and endpoint of } e_i \text{ in } V\backslash S \\ 0 & \text{otherwise} \end{cases}$$

And so, $f(S) = \sum_{i=1}^{n} f_{e_i}(S)$. Suppose (u, v) is an arbitrary edge in E, also suppose A, B are some arbitrary subsets of V. We wish to show in all cases, however vertices u and v appear or do not appear in A and/or B, that $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$:

<u>Case 1: $u \notin A$ and $u \notin B$:</u> Clearly, $u \notin (A \cup B)$ and $u \notin (A \cap B)$, thus, $f_{(u,v)}(A) = f_{(u,v)}(B) = f_{(u,v)}(A \cup B) = f_{(u,v)}(A \cap B) = 0$. So, $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds.

<u>Case 2: $u \in A$ and $u \notin B$ (W.L.O.G. swtiching B for A in this case and all subcases the claim holds):</u>

<u>Case 2.1: $v \in A$:</u> Clearly, $u, v \in (A \cup B)$ and $u \notin (A \cap B)$, it follows that $f_{(u,v)}(A) = f_{(u,v)}(B) = f_{(u,v)}(A \cup B) = f_{(u,v)}(A \cap B) = 0$. So, $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds.

<u>Case 2.2: $v \notin A$ and $v \in B$:</u> So, $f_{(u,v)}(A) = 1$ and $f_{(u,v)}(B) = 0$, but since $u, v \in (A \cup B)$ and $u \notin (A \cap B)$, then $f_{(u,v)}(A \cup B) = f_{(u,v)}(A \cap B) = 0$. So, $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds.

<u>Case 2.3: $v \notin A$ and $v \notin B$:</u> Again, $f_{(u,v)}(A) = 1$ and $f_{(u,v)}(B) = 0$. Also, $u \in (A \cup B)$ and $v \notin (A \cup B)$, then $f_{(u,v)}(A \cup B) = 1$. Again, $u \notin (A \cap B)$, then $f_{(u,v)}(A \cap B) = 0$. So, $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds.

<u>Case 3: $u \in A$ and $u \in B$ (W.L.O.G. swtiching B for A in this case and all subcases the claim holds):</u>

<u>Case 3.1: $v \in A$ and $v \in B$:</u> Since $u, v \in A$ and B, then $f_{(u,v)}(A) = f_{(u,v)}(B) = 0$. Also, $u, v \in (A \cup B)$ and $(A \cap B)$, then $f_{(u,v)}(A \cup B) = f_{(u,v)}(A \cap B) = 0$. So, $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds.

<u>Case 3.2: $v \notin A$ and $v \in B$:</u> It follows that $f_{(u,v)}(A) = 1$ and $f_{(u,v)}(B) = 0$. Since $u, v \in (A \cup B)$ then $f_{(u,v)}(A \cup B) = 0$, but $u \in (A \cap B)$ and $v \notin (A \cap B)$, thus $f_{(u,v)}(A \cap B) = 1$. Again, $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds.

<u>Case 3.3: $v \notin A$ and $v \notin B$:</u> Clearly, $f_{(u,v)}(A) = 1$ and $f_{(u,v)}(B) = 1$. Also, $u \in (A \cup B)$ and $(A \cap B)$, but $v \notin (A \cup B)$ and $(A \cap B)$, then $f_{(u,v)}(A \cup B) = f_{(u,v)}(A \cap B) = 1$. Then, $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds.

In all cases in which the vertices of some edge $(u, v) \in E$ might appear in $A, B \subseteq V$ it has been shown that $f_{(u,v)}(A) + f_{(u,v)}(B) \geqslant f_{(u,v)}(A \cup B) + f_{(u,v)}(A \cap B)$ holds. It must follow then that $\sum_{i=1}^{n} f_{e_i}(A) + \sum_{i=1}^{n} f_{e_i}(B) \geqslant \sum_{i=1}^{n} f_{e_i}(A \cup B) + \sum_{i=1}^{n} f_{e_i}(A \cap B) \rightarrow f(A) + f(B) \geqslant f(A \cup B) + f(A \cap B)$. Clearly, function f is submodular by the equivalence of submodular functions. ■

Q3. Consider the restructuring of LP formulation of maximum flow into standard form:

maximize: $\vec{c} \cdot \vec{x}$

subject to: A$\vec{x} \leqslant \vec{b}$, $\vec{x} \geqslant 0$.

Suppose this LP operates on G(V, E) such that paths P = {p_1, ..., p_n} and E = {e_1, ..., e_m}.

Let $\vec{c}$ = [1, ..., 1], such that $|\vec{c}|$ = n.

Let $\vec{x}$ = [x_1, ..., x_n], such that x_i = f_{p_i}, maximum flow along path p_i.

Let A be and m x n matrix and let A be defined as follows:

$$A[i,j] = \begin{cases} 1 & \text{edge } e_i \text{ appears on path } p_j \\ 0 & \text{edge } e_i \text{ does not appears on path } p_j \end{cases}$$

Let $\vec{b}$ = [c(e_1), ..., c(e_m)], such that $|\vec{b}|$ = m.

Using the restructured LP, we construct the dual as follows:

minimize: $\vec{b} \cdot \vec{y}$

subject to: $A^t\vec{y} \geqslant \vec{c}$, $\vec{y} \geqslant 0$.

Let $\vec{y}$ = [y_1, ..., y_m], y_i is a value associated with e_i.

Again, we define $\vec{c}$ and $\vec{b}$ as the same in the original construction.

Now, A^t is an n x m matrix, and so A^t is defined as follows:

$$A^t[i,j] = \begin{cases} 1 & \text{in path } p_i \text{ the edge } e_j \text{ appears} \\ 0 & \text{in path } p_i \text{ the edge } e_j \text{ does not appear} \end{cases}$$

The conclusion for the dual LP is that the objective function and constraints represent a construction of a max-flow min-cut, such that when the dual LP is complete its solution will equal the max-flow calculated by the original LP. It can now be stated:

$$y_i = \begin{cases} 1 & \text{edge } e_i \text{ connects sets S and T*} \\ 0 & \text{edge } e_i \text{ does not connect sets S and T*} \end{cases}$$

*clearly S and T represent the set of vertices separated by the minimum cut of G(V, E).

Then the objective function represents the summation of the capacity of those edges where y_i = 1, and the summation will be the maximum flow in the network G(V, E). The n path constraints, $A^t\vec{y} \geqslant \vec{c}$, represent the fact that on each path at least one edge connects S to T. Thus, for path p_i, $\sum_{j=1}^{m} a_{ij}y_j \geqslant 1$.

Q4. Polytope P is represented by $Ax \leqslant b$. Let $x \in \mathbb{R}^n$, so x is a point within P if $Ax \leqslant b$ holds. More specifically, let A be an m x n matrix, and b a series of m constants, such that $a_{i1}x_1 + ... + a_{in}x_n \leqslant b_i$ represents a closed half-space for each facet of the polytope, so that when all constraints are combined they represent all internal and boundary points of P.*

The LP must be designed so that we find a centre point y of a ball in P, such that all points on the surface of the ball that intersect in the plane where P exists are also in P. Suppose some point x exists at the border of P, then for some i, specifically the i that represents the facet on which x exists, $a_{i1}x_1 + ... + a_{in}x_n = b_i$. So, if the ball is the largest possible ball to fit in P the surface points must be as close to the boundaries as possible.

Suppose that $\vec{a}_i = [a_{i1}, ..., a_{in}]$ is a unit vector for some i, such that $\|\vec{a}_i\| = 1$. Indeed, $\vec{a}_i$ is the gradient unit vector to the facet represented by $a_{i1}x_1 + ... + a_{in}x_n = b_i$, it is also perpendicular to the facet. Also, as a gradient vector, $\vec{a}_i$ points in the direction outside the closed half-space defined by $a_{i1}x_1 + ... + a_{in}x_n \leqslant b_i$. To discover if the ball at center y with radius $r \in \mathbb{R}$ can fit in P we check for each facet i if the surface point $z_i = (\vec{a}_i * r + y)$, perpendicular to the facet i, is within P. To do this we use the constraint $a_{i1}z_{i1} + ... + a_{in}z_{in} \leqslant b_i$. To explain z_i further, $\vec{a}_i$*r is an r length vector that when added to y, projects a point to the surface or boundary of the ball and is perpendicular to facet i. In order to find the largest such ball, the LP must be designed to calculate both the ball's maximum radius r and its centre point y. This is done by establishing z_i's such that for all i, given constraints $a_{i1}z_{i1} + ... + a_{in}z_{in} \leqslant b_i$, we want $a_{i1}z_{i1} + ... + a_{in}z_{in}$ to be as close as possible to b_i for some maximum r and centre point y.

Thus, the LP is:
maximize:
r
subject to:
$\vec{a}_i \cdot (\vec{a}_i * r + y) \leqslant b_i, \forall\ i, 1 \leqslant i \leqslant m$
$r \geqslant 0$.

*explanation of polotope taken from wikipedia, http:\\en.wikipedia.org\wiki\Convex_polytope.

Q5. Suppose that P_1 and P_2 were to intersect, a primal LP that might define this intersection could be:
maximize: cx
subject to: $A_1x + A_2x \leqslant b_1 + b_2$

Also its dual LP:
minimize: $sb_1 + tb_2$
subject to: $sA_1 + tA_2 = c$, s, t $\geqslant 0$.

However, we know that P_1 and P_2 do not intersect, so in the primal LP no x will be found with the provided constraint. The primal LP is infeasible. In the dual LP suppose we let c be the zero vector, then clearly there exists an s and t such that $sA_1 + tA_2 = \vec{0}$. By the theorem of duality, where the primal LP has an optimal solution iff the dual has an optimal solution, it follows that our dual LP cannot have an optimal solution. If the dual had a feasible solution it would contradict the infeasibility of the primal LP.

If both LP's had optimal solutions then by weak duality we would have, $(sA_1 + tA_2)x = cx \leqslant sb_1 + tb_2$. Indeed, in the feasible case, the result of the dual LP would be an upperbound to the result of the primal LP. As stated previously, $c = \vec{0}$, and any point y dot product with c equals 0. It should be noted that the inequality $sb_1 + tb_2 \geqslant 0$ suggests that the dual has a minimum value, which if it did the result of the dual LP would be feasible, but this cannot be. In the situation where P_1 and P_2 do not intersect the objective function of the dual, $sb_1 + tb_2$, is unbounded. It follows that there exists s, t $\geqslant 0$ where $sA_1 + tA_2 = \vec{0}$ but $sb_1 + tb_2 < 0$. ■

Q6.

Q7.

Q8.