RECAP

Weighted Vertex Cover $\langle G, w \rangle$

- determine VC V^* of minimum weight $\sum_{v \in V^*} w(v)$
  - FAIR PRICING extends the maximal matching algorithm (unweighted case)
  - LP rounding:

$$\begin{cases} \min \sum_{v \in V} w(v) x_v \\ \text{s.t.} \\ x_u + x_v \geq 1 \; : \; \{u,v\} \in E \\ 0 \leq x_u \leq 1 \; : \; u \in V \end{cases}$$

$$V' = \{ v \in V : x_v^* \geq \tfrac{1}{2} \}$$

CURIOSITY!

DUAL:

(Ignore constraints $x_u \leq 1$)

REDUNDANT

$$\begin{cases} \max \sum_{e \in E} p_e \\ \text{s.t.} \\ \sum_{e \ni u} p_e \leq w(u) \quad u \in V \\ p_e \geq 0 \quad e \in E \end{cases}$$

Linear relaxation of fair pricing!

THE TRAVELLING SALESMAN PROBLEM (TSP)

SALESPERSON

Decision version:

TSP

I: $\langle G=(V,E), c, K \rangle$:

G complete, undirected graph

$c : V \times V \to \mathbb{N}$, symmetric

$(c(u,v) = c(v,u))$

$K \in \mathbb{N} - \{0\}$ $(c(u,u) = 0)$

Q: Is there a tour (hamiltonian cycle) $\gamma = \langle v_1, v_2, \ldots, v_{|V|}, v_{|V|+1} = v_1 \rangle$ $(v_i \neq v_j \;\; 1 \le i < j \le |V|)$ of cost

$$c(\gamma) = \sum_{i=1}^{|V|} c(v_i, v_{i+1}) \le K\ ?$$

OBSERVATION: There are $\sim |V|!$ tours in G

We can prove TSP $\in$ NPH by reducing from HAMILTON:

An istance of HAMILTON is an arbitrary undirected graph $G=(V,E)$.

Reduction:

$$\langle G=(V,E) \rangle \xrightarrow{f} \langle G'=(V,E'), c, K \rangle$$

with $E' = \{ \{u,v\} : u,v \in V \}$; (complete set)

$$c(u,v) = \begin{cases} 1 & \text{if } \{u,v\} \in E \\ 2 & \text{otherwise} \end{cases} ; \quad K = |V|$$

If $\langle G\rangle \in$ HAMILTON, then the cost of a hamiltonian cycle of G under $c(\cdot)$ is exactly $|V| \Rightarrow f(\langle G\rangle) = \langle G', c, |V|\rangle \in TSP$

If $f(\langle G\rangle) = \langle G, c, |V|\rangle \in TSP$ then there is a tour γ (made of $|V|$ edges) of cost $c(\gamma) = |V|$

$\Rightarrow$ ALL edges in the tour must cost 1

$\Rightarrow$ ALL edges in the tour belong to E

$\Rightarrow$ the tour is a hamiltonian cycle of G

$\Rightarrow$ $\langle G\rangle \in$ HAMILTON

Two interesting results on TSP:

1. Strong inapproximability result of general TSP:
   if $P \neq NP$ there cannot exist a poly-time $\rho(n)$-approximation for many functions $\rho(n)$

2. Good approximation algorithms for a "natural" (still hard) restriction of TSP (metric cost function)

   2-approximation (based on MST)

   $\frac{3}{2}$-approximation (Christofides' algorithm) (based on MST and min-cost matching)

INAPPROXIMABILITY RESULT FOR GENERAL TSP

THEOREM If $P \neq NP$ there cannot exist a polynomial $\rho(|V|)$-approximation algorithm for TSP, for ANY polynomial-time computable function $\rho(|V|)$.

Very strong result: observe that even $\rho(|V|) = 2^{|V|}$ is poly-time computable (generate 1 followed by $|V|$ 0's)

STANDARD TECHNIQUE: by contradiction: I show that if a $\rho(|V|)$-approximation algorithm exists, I can use it to solve an NPC problem in polynomial time (against the hypothesis $P \neq NP$)

PROOF We show that if A^{ρ}_{TSP} exists, I can solve HAMILTON in polynomial time;

Let $\langle G=(V,E) \rangle$ be an instance of HAMILTON. I create $\langle G' = (V, E'), c: V \times V \to \mathbb{N} \rangle$ an instance of the optimization TSP problem following the idea of the reduction above:

$$E' = \{\{u,v\} : u \neq v \in V\}; \quad c(u,v) = \begin{cases} 1 & \text{if } \{u,v\} \in E \\ |V| \cdot \rho(|V|) + 1 & \text{otherwise} \end{cases}$$

Observe that $\langle G', c \rangle$ can be computed in polynomial time since $\rho(|V|)$ is ptc.
Thus, $|\langle G', c\rangle| = poly(|\langle G \rangle|)$

Let us see what happens if I run $A^{\rho}_{TSP}(G', c)$:

1) If $\langle G \rangle \in$ HAMILTON, then G' contains a tour γ^* of cost $c(\gamma) = |V|$. But then $A^{\rho}_{TSP}(G')$ has to return a tour γ_A of cost $c(\gamma_A) \leq \rho(|V|) \cdot |V|$:
   $\Rightarrow \gamma_A$ cannot use any edge $\notin E$!
   $\Rightarrow \gamma_A$ is a hamiltonian cycle of G of cost $c(\gamma_A) = c(\gamma^*) = |V|$

2) If $\langle G \rangle \notin$ HAMILTON, then any tour γ of G must use an edge $e \notin E \Rightarrow \forall \gamma : c(\gamma) > |V| \cdot \rho(|V|)$

1) + 2) imply that
$\langle G \rangle \in$ HAMILTON $\Leftrightarrow c(\gamma_A = A^{\rho}_{TSP}(G', c)) \leq |V| \rho(|V|)$

Then:

```
DECIDE_HAMILTON (<G=(V,E)>)
  * build <G'=(V,E'), c> *
  γA ← A^ρ_TSP(G', c)
  if (c(γA) ≤ |V| ρ(|V|))
     then return 1
     else return 0
```

is a poly-time decision algorithm for HAMILTON, which contradicts $P \neq NP$

The proof implements the general GAP technique to prove inapproximability results:

GIVEN: minimization problem Π_m for which I want to prove $\rho(n)$-inapproximability:

DETERMINE: - decision problem $\Pi_d \in NPC$
- ptc function $f(x)$ transforming instances of Π_d into instances of Π_m such that:
  - Let $c: S \to \mathbb{N}^+$ be the cost function of Π_m: then exist a function $k(n)$ such that
    1. $x \in L_{\Pi_d} \Rightarrow c(s^*(f(x)) \leq k(|f(x)|)$
    2. $x \notin L_{\Pi_d} \Rightarrow c(s^*(f(x)) > \rho(|f(x)|) \cdot k(|f(x)|)$
- Then if $P \neq NP$ there cannot be a poly-time $\rho(n)$-approximation algorithm for Π_m

PROOF If $A^{\rho}_{\Pi_m}$ existed:

DECIDE(x)
  $s \leftarrow A^{\rho}_{\Pi_m}(A_f(x))$
  if $c(s) \leq \rho(|f(x)|) \cdot k(f(x))$
    then return 1
    else return 0

would decide Π_d in polynomial time

RESTRICTION: METRIC COST FUNCTION

DEF: A symmetric function $c: V \times V \to \mathbb{N}$ is metric if it satisfies the triangle inequality: $\forall u, v, w: c(u,v) \leq c(u,w) + c(w,v)$

Triangle ($\triangle$) inequality states that it is always more convenient to go directly from u to v rather than via intermediate nodes. E.g.: nodes u, v are points in $\mathbb{R}^d$ and $c(u,v)$ is their Euclidean distance

The decision problem TRIANGLE-TSP is defined as TSP: the instance is $\langle G, c, K \rangle$ with the additional constraint that c is metric

THEOREM TRIANGLE-TSP $\in$ NPH

PROOF We reduce from TSP as follows:

$$\langle G=(V,E), c, K \rangle \xrightarrow{f} \langle G=(V,E), c', K' \rangle$$

Let $W = \max\{c(u,v) : u, v \in V\}$:

$$c'(u,v) = \begin{cases} c(u,v) + W, & u \neq v \in V \\ 0 & u = v \in V \end{cases}$$

$$K' = K + |V| \cdot W$$

The reduction is correct, since c' satisfies $\triangle$:

$$\forall u \neq v \in V: c'(u,v) = c(u,v) + W \leq W + W$$
$$\leq c(u,w) + W + c(w,v) + W$$
$$= c'(u,w) + c'(w,v)$$

Clearly, f is ptc

$\langle G, c, K \rangle \in TSP \Leftrightarrow \exists$ tour γ in G with $c(\gamma) \leq K$

$\Leftrightarrow$ (since γ uses $|V|$ edges)

$\exists$ tour γ in G with $c'(\gamma) \leq \overbrace{K + |V|W}^{K'}$

$\Leftrightarrow \langle G, c', K' \rangle \in$ TRIANGLE-TSP

A 2-APPROXIMATION ALGORITHM FOR $\triangle$-TSP

We approximated VC by relating $|V^*|$ to $|A|$ a (maximal) matching of G. For $\triangle$-TSP we use a similar approach and relate the optimal tour γ^* to a Minimum Spanning Tree (MST) of G.

DEF: Given an undirected, connected, weighted $G = (V, E)$, $c: V \times V \to \mathbb{N}^+$, a Spanning Tree $T = (V, E_T)$ is any connected subgraph of G, with $|E_T| = |V| - 1$. We set $c(T) = \sum_{\{u,v\} \in E_T} c(u,v)$

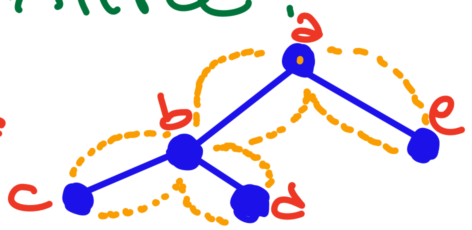
FACT A Minimum Spanning Tree (MST) of G can be determined in time $O(|E| \log |V|)$ (e.g. Prim's algorithm, 572 CLRS)

IDEA: I determine an MST T^* of G and obtain a TOUR γ_{T^*} from it.

HOW: γ_{T^*} is obtained as the preorder visit of T^* (closed on the first node)

EXAMPLE:

$T^* =$

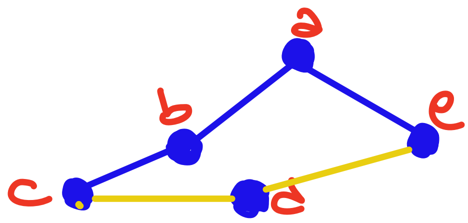


$\longrightarrow$ PREORDER VISIT: $\langle a, b, c, d, e \rangle$

$\downarrow$

TOUR: $\langle a, b, c, d, e, a \rangle$

$\gamma_{T^*} =$



We will prove that $c(\gamma_{T^*}) \leq 2c(\gamma^*)$!

Here is the algorithm:

```
APPROX_T_TSP(G, c)
  * V = {v_1, v_2, ---, v_|V|} *
  T* = (V, E_T) <- PRIM(G, c)
  r = v_1
  γ_T* <- <PREORDER(T*, r), r>
  return γ_T*
```

```
PREORDER(T, r)
P <- <r>
if (not leaf(r))
  then
    for each (s ∈ child(r))
      do
        P <- <P, PREORDER(T, s)>
return P
```

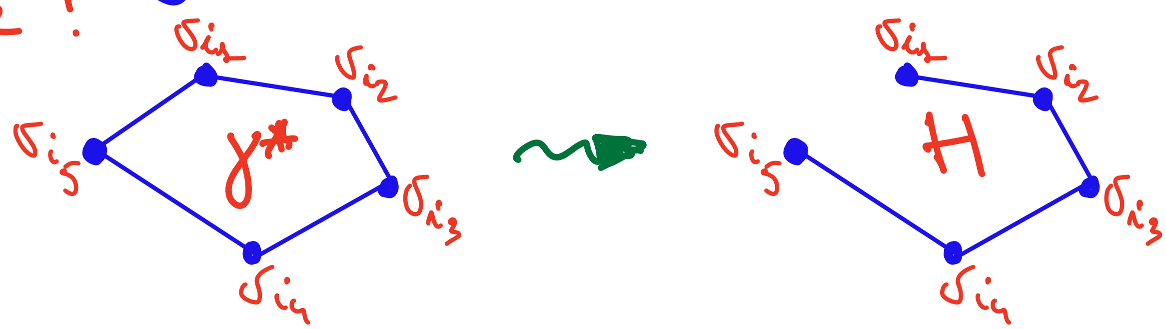
$$T_{APP}(|\langle G, c\rangle|) = \underbrace{O(|E|\log|V|)}_{T_{PRIM}} = O(|\langle G, c\rangle| \log |\langle G, c\rangle|)$$

To prove that A-T-T is a 2-approximation, we must relate OPT $= c(\gamma^*)$ and $c(T^*)$

LEMMA $c(\gamma^*) \geq c(T^*)$

PROOF Let $\gamma^* = \langle v_{i_1}, v_{i_2}, \dots, v_{i_{|V|}}, v_{i_1} \rangle$.

Remove edge $\{v_{i_{|V|}}, v_{i_1}\}$: we obtain a hamiltonian path $H = \langle v_{i_1}, v_{i_2}, \dots, v_{i_{|V|}} \rangle$ touching all nodes. H is a spanning Tree!



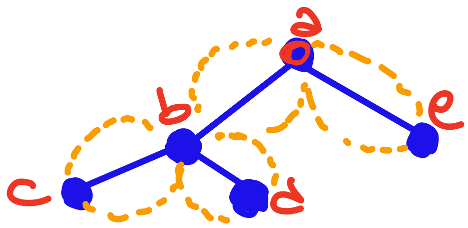
We have $\underset{\text{MST}}{c(T^*)} \leq \underset{\text{ST}}{c(H)} \leq c(\gamma^*)$!

The cost $c(T^*)$ is a lower bound to $c(\gamma^*)$ BUT:

observe that γ_{T^*} uses edges outside T^*! We have to upper bound $c(\gamma_{T^*})$ as a function of $c(T^*)$

DEF A Full Preorder Walk (FPW) is the nonsimple cycle obtained by considering the sequence of active calls of procedure PREORDER(T, r)

EXAMPLE:

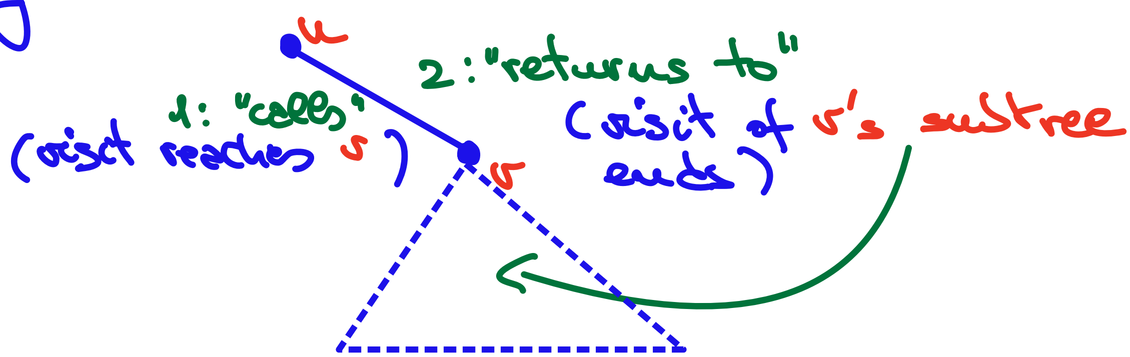


PREORDER(T, a) calls
PREORDER(T, b) calls
PREORDER(T, c) returns to
PREORDER(T, b) calls
PREORDER(T, d) returns to
PREORDER(T, b) returns to
PREORDER(T, a) calls
PREORDER(T, e) returns to PREORDER(T, a)

FPW = ⟨a, b, c, b, d, b, a, e, a⟩

PROPERTIES OF A FPW

1. The first occurrence of each node gives the preorder visit of T
2. Each edge of T is used $\boxed{\text{2 times}}$ by FPW:



Therefore: $\boxed{c(FPW) = 2c(T^*)}$