Convex Optimization

Spectral Graph Theory

This part is organized as follows: the first two sections discuss graph Laplacian in the spectral domain, by bounding its eigenvalues and introducing its importance (Cheeger's inequality). The random walk section provides a use case of graph spectral in the analysis of convergence, and introduces the importance of Laplacian linear equations. The next two sections show how to solve it exactly by applying pseudo inverse, as the Laplacian is non-invertible. The rest sections show how to approximate a graph Laplacian and solve the Laplacian linear equation approximately and more efficiently.

1 Spectral Domain of the Graph Laplacian

We assume the graph has n vertices and m edges. The vertex set is V and the edge set is E.

1.1 Courant-Fischer Theorem

1. **Eigenvalue version**: Let A be a symmetric matrix with eigenvalues $\lambda_1 \le \cdots \le \lambda_n$, then
$$\lambda_i = \min_{\substack{\text{subspace } W\subseteq\mathbb{R}^n\\ \dim(W)=i}} \max_{\substack{x\in W\\ x\neq\mathbf{0}}} \frac{x^\top Ax}{x^\top x} = \max_{\substack{\text{subspace } W\subseteq\mathbb{R}^n\\ \dim(W)=n+1-i}} \min_{\substack{x\in W\\ x\neq\mathbf{0}}} \frac{x^\top Ax}{x^\top x}.$$
2. **Eigenbasis version**: Let A be a symmetric matrix with eigenvalues $\lambda_1 \le \cdots \le \lambda_n$ and corresponding orthonormal eigenvectors $x_1,\ldots,x_n$, then
$$\lambda_i = \min_{\substack{x\perp x_1,\ldots x_{i-1}\\ x\neq\mathbf{0}}} \frac{x^\top Ax}{x^\top x} = \max_{\substack{x\perp x_{i+1},\ldots x_n\\ x\neq\mathbf{0}}} \frac{x^\top Ax}{x^\top x}.$$

Note that we also have $\lambda_i = \frac{x_i^\top Ax_i}{x_i^\top x_i}$.

Applying Courant-Fischer theorem, we have $\lambda_2 x^\top x \le x^\top Lx \le \lambda_n x^\top x$ for all $x\perp\mathbf{1}$, as $\mathbf{1}$ is the eigenvector of L corresponding to eigenvalue 0. For connected graphs, $\lambda_2 > 0$.

1.2 PSD Order (Loewner Order)

Defined only for symmetric matrices: $A \preceq B$ iff for all $x\in\mathbb{R}^n$, we have $x^\top Ax \le x^\top Bx$. We also define $G \preceq H$ for two graphs G and H iff $L_G \preceq L_H$. We always have $G \succeq H$ if H is a subgraph of G.

Properties:

1. If $A\preceq B$ and $B\preceq C$, then $A\preceq C$.
2. If $A\preceq B$, then $A+C\preceq B+C$ for any symmetric C.
3. If $A\preceq B$ and $C\preceq D$, then $A+C\preceq B+D$.
4. If $A\succ 0$ and $\alpha\ge 1$, then $\frac{1}{\alpha}A\preceq A\preceq \alpha A$.
5. If $A\preceq B$, then $\lambda_i(A)\le\lambda_i(B)$ for all i. Proof by Courant-Fischer theorem. The converse is not true.
6. For any matrix C, if $A\preceq B$, then $C^\top AC\preceq C^\top BC$.
7. If $0\preceq A\preceq B$, then $B^{-1}\preceq A^{-1}$.

1.3 Bounding the λ_2 and λ_n

1.3.1 Test Vector Method

Since $\lambda_2 \le \frac{y^\top Ly}{y^\top y}$ for any $y\perp\mathbf{1}$, we can upper bound the λ_2 by any **test vector** y. Similarly, we can lower bound the λ_n by test vectors by $\lambda_n \ge \frac{y^\top Ly}{y^\top y}$.

1. For a complete graph K_n, $L = n\boldsymbol{I} - \mathbf{1}\mathbf{1}^\top$ and for any $x\perp\mathbf{1}$ we have $x^\top Lx = nx^\top x$. Therefore, $\lambda_2(K_n) = \cdots = \lambda_n(K_n) = n$ and any $x\perp\mathbf{1}$ is an eigenvector.
2. For a path graph P_n, let $x(i) = n+1-2i$ be the test vector which satisfies $x\perp\mathbf{1}$, we get $\lambda_2(P_n)\le\frac{12}{n^2}$. Let $x(1)=-1$, $x(n)=1$ and $x(i)=0$ for other i to be the test vector, we get $\lambda_n(P_n)\ge 1$.
3. For a complete binary tree T_n (depth equals zero for a single root), let $x(i)=0$ for all non-leaf nodes, $x(i)=-1$ for even-numbered leaf nodes and $x(i)=1$ for odd-numbered leaf nodes be the test vector, we get $\lambda_n(T_n)\ge 1$. Let $x(1)=0$, $x(i)=1$ for the left subtree of the root and $x(i)=-1$ for the right subtree of the root be the test vector, we get $\lambda_2(T_n)\le\frac{2}{n-1}$.

1.3.2 Consequences of PSD Order

Since $x^\top(D-A)x = \sum_{(u,v)} w(u,v)(x(u)-x(v))^2 \ge 0$ and $x^\top(D+A)x = \sum_{(u,v)} w(u,v)(x(u)+x(v))^2 \ge 0$, we have $D\succeq A$ and $D\succeq -A$. In addition, we have $D\preceq(\max D_{i,i})\boldsymbol{I}$. Therefore, we have $L = D-A\preceq 2D\preceq(2\max D_{i,i})\boldsymbol{I}$, which implies $\lambda_n\le 2\max D_{i,i}$ for any graph. For unit-weight graphs, this means $\lambda_n\le 2\max\text{degree}(v)$. The bound is tight for a single-edge graph.

To get lower bounds of $\lambda_2(H)$, we first establish $f(n)H\succeq G$ for some G with known lower bounds on $\lambda_2(G)$. Usually $G=K_n$ because $\lambda_2(K_n)=n$. Then it follows that $\lambda_2(H)\ge\lambda_2(G)/f(n)$.

1. **Path Graph** P_n: Let $G_{i,j}$ denote a unit-weight graph consisting of one edge (i,j) and P_n be the path graph connecting 1 and n. Then $(n-1)P_n\succeq G_{1,n}$. Proof follows from applying Cauchy-Schwartz for $\delta_i := x(i+1)-x(i)$. For weighted paths, we have $G_{1,n}\preceq\left(\sum_{i=1}^{n-1}\frac{1}{w_i}\right)\sum_{i=1}^{n-1} w_iG_{i,i+1}$.
Applying path inequality, we have $K_n=\sum_{i<j}G_{i,j}\preceq\sum_{i<j}(j-i)P_{i,j}\preceq\sum_{i<j}(j-i)P_n\preceq n^3P_n$, which implies $\lambda_2(P_n)\ge\lambda_2(K_n)/n^3 = 1/n^2$.
2. **Any unit-weight graph** G: Define the diameter D of a graph G to be the maximum length of the shortest paths between any two nodes. Let $G_{i,j}^s$ be the shortest path from i to j. Applying path inequality, we have $K_n=\sum_{i<j}G_{i,j}\preceq\sum_{i<j}DG_{i,j}^s\preceq\sum_{i<j}DG\preceq n^2DG$, which implies $\lambda_2(G)\ge\frac{1}{nD}$.
3. **Complete Binary Tree** T_n: Define G_e be the single-edge graph with edge e, and $T_{i,j}$ be the unique path between i and j. Applying the weighted path inequality, we have $K_n=\sum_{i<j}G_{i,j}\preceq\sum_{i<j}\left(\left(\sum_{e\in T^{i,j}}\frac{1}{w(e)}\right)\left(\sum_{e\in T^{i,j}}w(e)G_e\right)\right)\preceq\left(\max_{i<j}\sum_{e\in T^{i,j}}\frac{1}{w(e)}\right)\left(\sum_{i<j}\sum_{e\in T^{i,j}}w(e)G_e\right)$. For e connecting level i and $i+1$ for $i\in[d-1]$, we set $w(e)=2^i$. Then $\max_{i<j}\sum_{e\in T^{i,j}}\frac{1}{w(e)}\le 4$. Since the number of occurrence of e in $T^{i,j}$ for any $i<j$ is upper bounded by $n^2 2^{-i}$, we have $\sum_{i<j}\sum_{e\in T^{i,j}}w(e)G_e\preceq\sum_e w(e)n^2 2^{-i}G_e=\sum_e n^2G_e=n^2T_n$. Therefore, $K_n\preceq 4n^2T_n$, which implies $\lambda_2(T_n)\ge\frac{1}{4n}$.

2 Conductance

Definitions:

1. **Conductance of a vertex subset**: Given $\emptyset\subset S\subset V$, the conductance $\phi(S):=\phi(S)=\frac{|E(S,V\setminus S)|}{\min\{\mathbf{vol}(S),\mathbf{vol}(V\setminus S)\}}$, where $\mathbf{vol}(S):=\sum_{v\in S}\text{degree}(v)$. Define $\mathbf{1}_S$ to be the n-dimensional vector with only 1 for the vertices of S and 0 for the vertices of $V\setminus S$. Assuming $\mathbf{vol}(S)\le\mathbf{vol}(V)/2$, thus $|E(V,V\setminus S)|=\sum_{(u,v)\in E}(\mathbf{1}_S(u)-\mathbf{1}_S(v))^2=\mathbf{1}_S^\top L\mathbf{1}_S$ and $\mathbf{vol}(S)=\mathbf{1}_S^\top D\mathbf{1}_S$. Then $\phi(S)=\frac{\mathbf{1}_S^\top L\mathbf{1}_S}{\mathbf{1}_S^\top D\mathbf{1}_S}$.
2. **Conductance of a graph**: The conductance $\phi(G):=\min_{\emptyset\subset S\subset V}\phi(S)=\min_{\substack{\emptyset\subset S\subset V\\ \mathbf{vol}(S)\le\mathbf{vol}(V)/2}}\phi(S)$.
3. **ϕ-expander**: For any $\phi\in(0,1]$, we call a graph G to be a ϕ-expander if $\phi(G)\ge\phi$.
4. **ϕ-expander decomposition of quality** q: A partition $\{X_i\}$ of the vertex set V is called a ϕ-expander decomposition of quality q if (1) each induced graph $G[X_i]$ is a ϕ-expander, and (2.i) #edges not contained in any $G[X_i]$ is at most $q\cdot\phi\cdot m$. The second condition is equivalent to (2.ii) The partition removes at most $q\cdot\phi\cdot m$ edges.
5. **Normalized Laplacian**: We define the *normalized Laplacian* to be $N:=D^{-1/2}LD^{-1/2}$. N is still PSD, with first eigenvalue equals 0 associated with eigenvector $D^{1/2}\mathbf{1}$. By Courant-Fischer theorem, $\lambda_2(N)=\min_{x\perp D^{1/2}\mathbf{1}}\frac{x^\top Nx}{x^\top x}=\min_{z\perp d}\frac{z^\top Lz}{z^\top Dz}$.

2.1 Cheeger's Inequality

Notice that the $\lambda_2(N)$ has similar forms to $\phi(G)$. Cheeger's inequality aims to bound $\phi(G)$ by $\lambda_2(N)$.

Cheeger's Inequality: $\frac{\lambda_2(N)}{2}\le\phi(G)\le\sqrt{2\lambda_2(N)}$.

The lower bound is proved by restricting the minimum in $\lambda_2(N)$ to be $z_S=\mathbf{1}_S-\alpha\mathbf{1}$ for some α such that $z_S\perp d$. The upper bound is proved by constructing S for any $z\perp d$ such that $\frac{\mathbf{1}_S^\top L\mathbf{1}_S}{\mathbf{1}_S^\top D\mathbf{1}_S}\le\sqrt{2\frac{z^\top Lz}{z^\top Dz}}$.

3 Random Walks on a Graph

A **random walk on a graph** G is a Markov Chain with transition probability $\mathbb{P}(v_{t+1}=v\mid v_t=u)=w(u,v)/d(u)$ iff $(u,v)\in E$ and 0 otherwise. The transition matrix is thus $W=AD^{-1}=I-D^{1/2}ND^{-1/2}$ and $p_t=W^tp_0$. Define $\boldsymbol{\pi}=\frac{d}{\mathbf{1}^\top d}$, thus $\boldsymbol{\pi}=W\boldsymbol{\pi}$ for any G, so every G has a stationary distribution.

3.1 Lazy Random Walks

A **lazy random walk on a graph** G is a random walk, but has half probability to not move for every step. Assuming that

G is connected, the lazy random walk guarantees ergodicity of the Markov Chain, and thus convergence to the stationary distribution. The transition matrix is $\tilde{W} = \frac{1}{2}(I + W) = I - \frac{1}{2}D^{1/2}ND^{-1/2}$.

Relation between lazy random walk and normalized Laplacian: For the i-th eigenvalue ν_i of N associated with eigenvector ψ_i, the $\tilde{W}$ has an eigenvalue $1 - \frac{1}{2}\nu_i$ associated with eigenvector $D^{1/2}\psi_i$. Since $0 \preceq L \preceq 2D$, we have $0 \preceq N \preceq 2\boldsymbol{I}$ and thus $0 \leq \lambda_i(N) \leq 2$. Therefore, we conclude that all eigenvalues of $\tilde{W} \in [0,1]$.

Dynamics of lazy random walk: Expanding the starting distribution p_0 by the eigenvectors of $\tilde{W}$, we have for some $\{\alpha_i\}$ that $p_0 = \sum_{i=1}^n \alpha_i D^{1/2}\psi_i$. Therefore, we have $p_t = \tilde{W}^t p_0 = \sum_{i=1}^n \alpha_i(1-\frac{1}{2}\nu_i)^t D^{1/2}\psi_i \to \alpha_1 D^{1/2}\psi_1$ as $\nu_1 = 0$ and $\nu_i > 0$ for $i \neq 1$. Since $\psi_1 \propto D^{1/2}\mathbf{1}$, we have $\psi_1 = \frac{d^{1/2}}{(\mathbf{1}^\top d)^{1/2}}$, thus $\alpha_1 = \psi_1^\top D^{-1/2}p_0 = \frac{\mathbf{1}^\top p_0}{(\mathbf{1}^\top d)^{1/2}} = \frac{1}{(\mathbf{1}^\top d)^{1/2}}$ and $\alpha_1 D^{1/2}\psi_1 = \pi$, which implies $p_t \to \pi$, the stationary distribution.

Rate of Convergence: For any unit-weight connected graph G and any starting distribution p_0, we have $\|p_t - \pi\|_\infty \leq e^{-\nu_2 t/2}\sqrt{n}$. Therefore, a larger ν_2 and smaller vertex set means faster convergence, and the convergence rate is exponential.

3.2 Hitting Time

The expected hitting time from a to s is defined by $\mathbb{E}H_{a,s}$, where $H_{a,s} = \mathbf{argmin}_t\{v_t = s \mid v_0 = a\}$. We want $\mathbb{E}H_{a,s}$ for all vertices a and denote the vector as h, *e.g.*, $h(s) = 0$.

By one-step analysis, we have $h(a) = 1 + \sum_{(a,b)\in E} \frac{w(a,b)}{d(a)} h(b) = 1 + \mathbf{1}_a^\top W^\top h$, and thus $1 = \mathbf{1}_a^\top(\boldsymbol{I} - W^\top)h$. Combining the equation for all vertices except s, we have $\mathbf{1} - \alpha\mathbf{1}_s = (\boldsymbol{I} - W^\top)h$, where α represents the extra freedom from the $n-1$ equations. Multiplying both side by D, we get $d - \alpha d(s)\mathbf{1}_s = (D - A)h = Lh$, which only have solution when $d - \alpha d(s)\mathbf{1}_s \perp \mathbf{1}$. Therefore, $\alpha = \|d\|_1/d(s)$.

To summarize, by solving $Lh = d - \|d\|_1\mathbf{1}$, we can get the expected hitting time from all vertices to s. Note that the solution has one extra freedom because $\dim(\ker(L)) = 1$, and the correct expected hitting time is $h - h(s)\mathbf{1}$ to enforce the constraint that $h(s) = 1$. The equation can be solved in $\tilde{O}(m)$.

4 Pseudo-Inverse and Effective Resistance

Given a Laplacian L, its (Moore-Penrose) pseudo inverse is defined to be either of the two equivalents:

1. A matrix L^+ that is (1) symmetric, (2) $L^+v = 0$ for $v \in \ker(L)$, and (3) $L^+Lv = LL^+v = v$ for $v \in \ker(L)$.
2. Let λ_i, v_i be the i-th eigenvalue and eigenvector. Then $L^+ = \sum_{\lambda_i \neq 0} \lambda_i^{-1} v_i v_i^\top$.

Property:

- Assume $M = XYX^\top$, where X is real and invertible, and Y is real and symmetric. Let Π_M be the orthogonal projection to the image of M. Then $M^+ = \Pi_M(X^\top)^{-1}Y^+X^{-1}\Pi_M$.
- For symmetric L, $\Pi_L := \sum_{\lambda_i \neq 0} v_i v_i^\top = L^{+/2}LL^{+/2} = L^+L = LL^+$ is the orthogonal projection to the image of L, *i.e.*, $\Pi_L v = 0$ for any $v \in \ker(L)$ and $\Pi_v = v$ for any $v \in \mathbf{im}(L)$. For connected G, $\Pi_{L_G} = \boldsymbol{I} - \frac{1}{n}\mathbf{1}\mathbf{1}^\top$.

The effective resistance between vertex a and b is defined to be the cost (energy lost) to routing one unit (of positive electric charge) from a to b: $R_{\text{eff}}(a,b) = \min_{Bf = \mathbf{1}_b - \mathbf{1}_a} f^\top R f = \tilde{f}^\top R \tilde{f}$, where $\tilde{f}$ is the electric flow. Let $\tilde{x}$ be the electric voltages, we also have $L\tilde{x} = \mathbf{1}_b - \mathbf{1}_a$, and thus $R_{\text{eff}}(a,b) = \tilde{x}^\top L \tilde{x} = (\mathbf{1}_b - \mathbf{1}_a)^\top L^+(\mathbf{1}_b - \mathbf{1}_a) = \|L^{+/2}(\mathbf{1}_b - \mathbf{1}_a)\|_2^2$.

Effective Resistance is a distance defined on the vertex pairs, i.e. $R_{\text{eff}}(a,c) \leq R_{\text{eff}}(a,b) + R_{\text{eff}}(b,c)$.

5 Solving Laplacian Linear Equations Exactly

5.1 Optimization View

Solving $Lx = d$ is equivalent to solving $\mathbf{argmin}_x -d^\top x + \frac{1}{2}x^\top Lx$. By iteratively optimize over x_i, we get a series of similar optimizations. The final optimization is straightforward, then we can back substitute to get x.

5.2 Additive View

Given an invertible square lower/upper triangular matrix M, we can solve $Mx = d$ by back substitution in $O(\mathbf{nnz}(M))$, where $\mathbf{nnz}(M)$ means the number of non-zeros in M. Therefore, if we know the **Cholesky decomposition** $L = MM^\top$ (requires $O(n^3)$), then we can solve $Lx = d = M(M^\top x)$ by (1) solving $My = d$ then (2) solving $M^\top x = y$ in $O(\mathbf{nnz}(M))$.

However, the Laplacian is non-invertible, leading to one diagonal of M equals 0. Therefore, we need to play a trick. Define $\hat{M}$ equals M but has value 1 for the zero diagonal, and $\hat{D}$ be a diagonal matrix that has value 0 at the zero diagonal of M and 1 otherwise. Then $L = \hat{M}\hat{D}\hat{M}^\top$, and each $\hat{M}$ is now invertible. Since $\hat{D}^+ = \hat{D}$, we can find a special solution of $Lx = d$ by (1) solving $\hat{M}z = d$, (2) computing $y = \hat{D}z$, and (3) solving $\hat{M}^\top x = y$. The solution space is obtained by adding a subspace spanned by $\mathbf{1}$.

6 Approximating a Dense Graph in the Spectral Domain

6.1 Concentration of Random Matrices

1. **Chernoff Bound for Bounded independent variables**: Suppose $\{X_i \in \mathbb{R}\}$ are independent random variables and $0 \leq X_i \leq R$. Let $X = \sum_i X_i$ and $\mu = \mathbb{E}X$. Then for any $0 < \epsilon \leq 1$, we have $\mathbb{P}(X \geq (1+\epsilon)\mu) \leq \exp(-\frac{\epsilon^2\mu}{4R})$ and $\mathbb{P}(X \leq (1-\epsilon)\mu) \leq \exp(-\frac{\epsilon^2\mu}{4R})$.
2. **Bernstein Bound for independent, zero-mean and bounded variables**: Suppose $\{X_i \in \mathbb{R}\}$ are independent, zero-mean random variables and $|X_i| \leq R$. Let $X = \sum_i X_i$, and $\sigma^2 = \mathbf{Var}(X)$. Then for any $t > 0$, we have $\mathbb{P}(|X| \geq t) \leq 2\exp(\frac{-t^2}{2Rt + 4\sigma^2})$.
   The proof is similar to Chernoff bound. (1) $\mathbb{P}(X \geq t) = \mathbb{P}(\exp(\theta X) \geq \exp(\theta t)) \leq \exp(-\theta t)\mathbb{E}(\exp(\theta X))$, (2) upper bound $\mathbb{E}(\exp(\theta X)) \leq \exp(\theta^2\sigma^2)$ given $\theta \in (0, \frac{1}{R}]$, which allows $\exp(\theta X_i) \leq 1 + \theta X_i + (\theta X_i)^2$, and (3) take the minimum among $\theta \in (0, \frac{1}{R}]$.
3. **Bernstein Bound for independent, zero-mean and bounded symmetric matrices**: Suppose $\{X_i \in \mathbb{R}^{n\times n}\}$ are independent, zero-mean, symmetric random matrices and $\|X_i\| \leq R$, where $\|\cdot\|$ is the spectral norm (the largest singular value). Let $X = \sum_i X_i$, and $\sigma^2 = \|\sum_{i=1}^n \mathbb{E}X_i^2\|$. Then $\mathbb{P}(\|X\| \geq t) \leq 2n\exp(\frac{-t^2}{2Rt + 4\sigma^2})$.

6.2 Matrix Functions

Given a real-valued function $f: \mathbb{R} \to \mathbb{R}$ and a symmetric matrix A with eigen-decomposition $A = V\Lambda V^\top$, we define $f(A) = Vf(\Lambda)V^\top$. This is compatible to the Taylor expansion $f(x) = \sum_i \alpha_i x^i$, as $f(A) = \sum_i \alpha A^i = V(\sum_i \alpha_i f(\Lambda))V^\top = Vf(\Lambda)V^\top$.

Monotonicity: Given $f: \mathcal{D} \to \mathcal{C}$ and partial orders $\leq_{\mathcal{C}}$ and $\leq_{\mathcal{D}}$, we call f is monotonically increasing w.r.t. these orders iff for all $d_1 \leq_{\mathcal{D}} d_2 \in \mathcal{D}$ we have $f(d_1) \leq f(d_2)$. For matrix functions, we use the PSD order as the ordering.

Property:

- If the scalar function f is monotonically increasing, then the matrix function $X \to \mathbf{Tr}(f(X))$ is monotonically increasing.
- $\log(\cdot)$ is monotonically increasing.
- The matrix function $(\cdot)^2$ and $\exp(\cdot)$ is **not** monotone.
- $\exp(A) \preceq \boldsymbol{I} + A + A^2$ for $\|A\| \leq 1$.
- $\log(\boldsymbol{I} + A) \preceq A$ for $A \succ -\boldsymbol{I}$.
- (Lieb's theorem): $f(A) := \mathbf{Tr}(\exp(H + \log(A)))$ for some symmetric H is concave in the domain of PSD matrices. This is in particular useful with Markov inequality because $\mathbb{P}(\|X\| \geq t) = \mathbb{P}(\lambda_n \geq t) \leq \mathbb{P}(\mathbf{Tr}(\exp(\theta X)) \geq \exp(\theta t)) \leq \exp(-\theta t)\mathbb{E}(\mathbf{Tr}(\exp(\theta X)))$.

6.3 Spectral Sparsifiers

Given PD matrices A, B and $\epsilon > 0$, we say $A \approx_\epsilon B$ iff $\frac{1}{1+\epsilon}A \preceq B \preceq (1+\epsilon)A$. If $L_G \approx_\epsilon L_{\tilde{G}}$ and $|\tilde{E}| \ll |E|$, we call $\tilde{G}$ a spectral sparsifier of G.

Properties:

- Define $c_G(T) := \sum_{e \in E \cap (T \times V \setminus T)} w(e)$ to be the value of the cut $(T, V \setminus T)$. If $L_G \approx_\epsilon L_{\tilde{G}}$, then for all $T \subset V$, we have $\frac{1}{1+\epsilon}c_G(T) \leq c_{\tilde{G}}(T) \leq (1+\epsilon)c_G(T)$. The proof is by noticing $c_G(T) = 1_T^\top L_G 1_T$.
- $L \approx_\epsilon \tilde{L} \Leftrightarrow \Pi_L \approx_\epsilon L^{+/2}\tilde{L}L^{+/2}$, as $A \preceq B$ implies $C^\top AC \preceq C^\top BC$ for any $C \in \mathbb{R}^{n\times n}$.
- For $\epsilon \leq 1$, if $\|\Pi_L - L^{+/2}\tilde{L}L^{+/2}\| \leq \epsilon/2$, then $\Pi_L \approx_\epsilon L^{+/2}\tilde{L}L^{+/2}$.

Theorem: Consider a connected graph $G = (V, E, w)$, with $n = |V|$. For any $0 < \epsilon < 1$ and $0 < \delta < 1$, there exist sampling probabilities p_e for each edge $e \in E$ s.t. if we include each edge e in $\tilde{E}$ independently with probabilty p_e and set its weight $\tilde{w}(e) = \frac{1}{p_e}w(e)$, then with probability at least $1 - \delta$ the graph

$\tilde{G} = (V, \tilde{E}, \tilde{w})$ satisfies $L_G \approx_\epsilon L_{\tilde{G}}$ and $|\tilde{E}| \le O\left(n\epsilon^{-2}\log(n/\delta)\right)$. The proof uses Bernstein bounds to prove the concentration of the constructed random graph.

7 Solving Laplacian Linear Equations Approximately

Idea: solving Laplacian linear equations requires $O(n^3)$ to get the Cholesky decomposition, which is expensive when the graph is large. By approximating the Laplacian, we can get an approximation of the solution quickly, especially in sparse graphs.

Given PSD matrix M and $d \in \mathbf{im}(M)$, let $Mx^* = d$. We say that $\tilde{x}$ is an ϵ-approximate solution to $Mx = d$ iff $\|\tilde{x} - x^*\|_M^2 \le \epsilon \|x^*\|_M^2$, where $\|x\|_M^2 = x^\top Mx$. Note that any solution to $Mx = d$ has the same $\|\cdot\|_M^2$, as they differ by a vector in the kernel of M.

Theorem: Given a Laplacian $\boldsymbol{L}$ of a weighted undirected graph $G = (V, E, \boldsymbol{w})$ with $|E| = m$ and $|V| = n$ and a demand vector $\boldsymbol{d} \in \mathbb{R}^V$, we can find $\tilde{\boldsymbol{x}}$ that is an ϵ-approximate solution to $\boldsymbol{L}\boldsymbol{x} = \boldsymbol{d}$, using an algorithm that takes time $O(m \log^c n \log(1/\epsilon))$ for some fixed constant c and succeeds with probability $1 - 1/n^{10}$. Note that without knowing the Cholesky decomposition in advance, the exact solution requires $O(n^3)$ and $m \le n^2/2$.

Combinatorial Graph Algorithms

8 Maximum Flows and Minimum Cuts

8.1 Flows

Definition:

1. s-t **flow**: an s-t flow is a flow such that $\boldsymbol{B}\boldsymbol{f} = F(-\mathbf{1}_s + \mathbf{1}_t)$ for some $F \ge 0$, *i.e.*, routes some unit from s to t. F is defined to be $\mathbf{val}(\boldsymbol{f})$.
2. **Maximum flow problem**: Given a source vertex and a sink vertex, maximize F such that $\boldsymbol{B}\boldsymbol{f} = F(-\mathbf{1}_s + \mathbf{1}_t)$ and $0 \le \boldsymbol{f} \le \boldsymbol{c}$.
3. **Path flow**: an s-t path flow is an s-t flow that only uses a simple path from s to t.
4. **Cycle flow**: A cycle flow is a flow that only uses a simple cycle, so it does not create net-in or net-out.

Path-cycle decomposition lemma: Any s-t flow can be decomposed to a sum of s-t path flows and cycle flows such that the summation has at most $\mathbf{nnz}(\boldsymbol{f})$ terms.

There is always an optimal flow that can be decomposed to only path flows, as the cycle flow does not route anything from s-t and removing all cycle flows in an optimal flow creates another optimal flow.

8.2 Cuts

Definition:

1. s-t **cuts**: an s-t cut is a cut $(S, V \setminus S)$ such that $s \in S$ and $t \in V \setminus S$.
2. **Minimum cut problem**: Given two vertices s and t, minimize the cut value $c_G(S) = \sum_{e \in E \cap (S \times V \setminus S)} w(e)$ such that $s \in S$ and $t \in V \setminus S$.

If there is no feasible s-t flow, then define S to be the set of vertices reachable from s, $(S, V \setminus S)$ is an s-t cut.

8.3 Solving Maximum Flow

Greedily adding flows on the original graph G leads to problems, but greedily adding flows on the residual graph G_f is optimal. This is because residual graph allows to cancel some part of the added flow in order to increase the unit routed.

The algorithm: (1) initialize $\boldsymbol{f} = 0$, (2) repeatly find an s-t flow $\tilde{\boldsymbol{f}}$ such that $-\boldsymbol{f} \le \tilde{\boldsymbol{f}} \le \boldsymbol{c} + \boldsymbol{f}$ and set $\boldsymbol{f} = \boldsymbol{f} + \tilde{\boldsymbol{f}}$.

Property:

1. Assume $\boldsymbol{f}$ is feasible in G. Then $\tilde{\boldsymbol{f}}$ is feasible in $G_f \Leftrightarrow \tilde{\boldsymbol{f}} + \boldsymbol{f}$ is feasible in G. Proof follows from definition.
2. A feasible $\boldsymbol{f}$ is optimal iff there is no feasible s-t flow in G_f. Proof by contradiction.

Ford-Fulkerson Algorithm

We call the minimum capacity of all edges in an s-t flow to be the bottleneck capacity.

Algorithm: find an arbitrage s-t path flow in G_f, augment it to route the bottleneck, then add it to the current flow, repeatly.

For irrational capacities this algorithm may not terminate. For integer capacities (rational capacities can be translated to integer capacities to multiplication), each round must increase the capacity of current flow by at least 1, so it terminates in F^*) augmentations, which is $O(mF^*)$) time.

Modified algorithm (may be faster in some cases): find the s-t path flow with the maximum bottleneck capacity, then add it to the current flow, repeatly.

Using binary search on the threshold of bottleneck capacity (only use edges with capacity greater than the threshold), we can find the maximum bottleneck capacity in $O(m \log n)$. This path flow carries at least $\frac{1}{m}$ fraction of the remaining flows in G_f, as there are at most m path flows. Therefore, it terminates when $(1 - \frac{1}{m})^T F^* < 1$, which means $T = O(m \log F^*)$. The total time is $O(Tm \log n) = O(m^2 \log n \log F^*)$.

8.4 Duality of Max Flow and Min Cut

- **Max flow $\le$ Min cut**: For any feasible s-t flow and any s-t cut, we have $\mathbf{val}(f) \le c_G(S)$. To see this, simply observe that this flow must cross the cut, so the maximum value that can be routed is bounded by the maximum capacity allowed by the cut. In particular, the maximum s-t flow is bounded by the minimum s-t cut.
- **Max flow $\ge$ Min cut**: let $\boldsymbol{f}$ be the maximum flow composed of only s-t path flows, then t is not reachable from s in G_f. Define S to be the vertex set that is reachable from s. Then $\boldsymbol{f}$ saturates every edge in $E \cap \{S \times V \setminus S\}$. There is no edge in $\boldsymbol{f}$ that is directed from $V \setminus S$ to S, as there is no edge from S to $V \setminus S$ in G_f. This implies $\mathbf{val}(\boldsymbol{f}) \ge c_G(S)$, and thus the maximum flow is greater than the minimum cut.

Combining this two, we estabilish that strong duality between the maximum flow and the minimum cut holds, *i.e.*, max flow = min cut.

9 Dinic's Algorithm for Maximum Flow